Dear TOS editors and reviewers,

Attached please find our submission to the ACM Transactions on Storage Systems.

Our paper is titled "*End-to-End Abstractions for Application-Aware Storage*." In this article, we provide an overview of the problem of "information-gap" in the storage stack, and present two novel abstractions that effectively bridge this gap, thereby enabling a range of functionality that is almost impossible to achieve with existing systems and interfaces. Most of the material in this article forms part of Gopalan Sivathanu's Ph.D. dissertation.

Our first abstraction is *Type-Aware Storage* that aims communicating pointer information to the disk hardware. We have published this abstraction in OSDI 2006. This article includes a new unpublished case-study of type-aware storage,"Disk-Level Data Consistency." This case-study proposes and evaluates how complex higher-level consistency properties can be achieved at the disk hardware-level, in a file-system–agnostic manner.

Our second abstraction is *Context-Aware I/O*, a flexible mechanism to communicate between applications and data, across the storage stack. We present the design, implementation, and evaluation of the above abstraction, and demonstrate its usefulness through two separate case-studies. This abstraction and its case-studies have not been published in any other venue.

Overall, of this 60 page article, about 60 percent is new unpublished material.

This work was completely done when all authors were affiliated with the File systems and Storage Laboratory at Stony Brook University, New York.

Thank you for your time and effort in reviewing this article.

Sincerely,

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End-to-End Abstractions for Application-Aware Storage

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Modern computer systems are a composition of several logically independent layers. Althought providing many important benefits, this rampant layering has also led to the well-explored problem of information-divide in the systems stack. Layers hide information, thus constraining functionality and limiting the power of individual layers. A particularly striking instance of this general problem exists in the storage stack today. Modern high-end storage systems have significant processing capabilities, but despite their potential, storage systems are constrained in their functionality because they are oblivious of higher layers and the applications using them. In this article, we seek to answer a simple question: *how can we convey application-level information across the diverse modern storage stack in a simple and generic manner?* We propose two flexible abstractions to solve this problem. The first abstraction is the notion of type-awareness in the storage stack. In type-aware storage, lower layers of the storage stack such as the disk are aware of the pointer relationships between disk blocks that are imposed by higher layers such as the file system. Type-awareness enables semantics-aware optimizations and new functionality in the lower layers of the storage stack. The second abstraction we describe is Context-Aware I/O (CAIO), a generic mechanism to propagate information end-to-end through the storage stack. CAIO provides a simple, yet effective interface to communicate *application-data* and *application-I/O* relationships to the storage stack, enabling interesting functionality. Through several case studies, we demonstrate the flexibility and benefits of both abstractions and show that they present a simple yet effective general interface to build the next generation of storage systems.

Categories and Subject Descriptors: D.4.2 [**Operating Systems**]: Storage Management—*Storage hierarchies;*

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D.4.2 [**Operating Systems**]: Storage Management—*Secondary storage*; D.4.2 [**Operating Systems**]: Storage Management—*Allocation/deallocation strategies*; D.4.2 [**Operating Systems**]: File Systems Management—*Access methods*; D.4.2 [**Operating Systems**]: File Systems Management—*File organization*; D.4.6 [**Operating Systems**]: Security and Privacy Protection—*Access controls*; D.4.6 [**Operating Systems**]: Security and Privacy Protection—*Authentication*; D.4.7 [**Operating Systems**]: Organization and Design—*Hierarchical Design*

General Terms: Design, Experimentation, Reliability, Security, Measurement, Performance

Additional Key Words and Phrases: Storage Stack, Intelligent Disks, File Systems, Storage Systems

1. INTRODUCTION

Computer system design over the past years has revolved around the principle of layering [Dijkstra 1968]. Building systems as a hierarchy of layers enables localized and independent innovation in the individual layers. For example, in the network protocol stack comprising layers such as application, transport, network, and data link—each layer can independently innovate as long as the interface exported to the other layers is intact. With the growing complexity of today's systems, layering has become and indispensable technique in hardware and software design.

Despite its obvious benefits, layered system design also comes with an inevitable side-effect: information available at one layer is not visible at the other layers beyond what is permitted by the interface separating those layers. The impact of this lack of information is becoming more pronounced in the recent years as there is a need for individual layers to support advanced functionality, requiring cross-layer information. This problem is exacerbated by the fact that recent advancements in computer systems such as virtual machine technology [Barham et al. 2003] have introduced more layers of virtualization in the systems stack, further widening this information-gap. Techniques to address this general problem of the information-gap across layers have ranged from building application-extensible OSes [Bershad et al. 1995; Engler et al. 1995] and brand-new abstractions [Mesnier et al. 2003; Sivathanu et al. 2006; MacCormick et al. 2004; Denehy et al. 2002], to more evolutionary approaches such as applications passing hints [Patterson et al. 1995; Cao et al. 1996; de Jonge et al. 2003], applications implicitly influencing OS behavior [Arpaci-Dusseau and Arpaci-Dusseau 2001; Burnett et al. 2002], and automatically inferring cross-layer information [Arpaci-Dusseau and Arpaci-Dusseau 2001; Sivathanu et al. 2003].

In the modern storage hierarchy, the general problem of information-gap between layers has hampered development of new functionality. Large-scale storage systems today comprise diverse resources that include high processing power, hundreds of gigabytes of RAM, solid state storage media such as flash, and hundreds or even thousands of disks [EMC Corporation 1999; Network Appliance Inc. 2006]. Despite these advancements in storage hardware, storage systems are constrained in the range of functionality they can provide, because they lack information about higher-level data semantics.

Techniques to address this general problem of the information-gap across layers have ranged from building application-extensible OSes [Bershad et al. 1995; Engler et al. 1995] and brand-new abstractions [Mesnier et al. 2003; Sivathanu et al. 2006; MacCormick et al. 2004; Denehy et al. 2002], to more evolutionary approaches such as applications passing hints [Patterson et al. 1995; Cao et al. 1996; de Jonge et al. 2003], applications implicitly

influencing OS behavior [Arpaci-Dusseau and Arpaci-Dusseau 2001; Burnett et al. 2002], and automatically inferring cross-layer information [Arpaci-Dusseau and Arpaci-Dusseau 2001; Sivathanu et al. 2003]. However, none of the existing solutions enable conveying both *application-data* and *application-I/O* relationships to the storage stack, in an end-to-end fashion (user applications to the storage hardware).

Our approach to solve the problem of information-gap is to propagate *minimal* and *generic* information relating to data and I/O, from higher-level layers of the storage stack to the lowest-level (the storage hardware). We developed two generic abstractions to encode *structural* and *operational* information available at the application-level and communicate it as part of I/O operations. Our first abstraction is *type-awareness*, which is to communicate **pointers** between disk blocks to the lower layers of the storage stack. Pointers establish relationships between disk blocks in a generic manner, and are maintained by layers such as file systems or databases. Our second abstraction is *context-aware storage*, which is to communicate higher-level **logical context** of I/O operations across the storage stack. For example, all I/O operations generated from a single user application can be grouped under the same logical context.

The following are the three key characteristics of our approach that differentiate our work from previous approaches:

(1) The information being communicated from higher-level layers is already available at the corresponding layers (e.g., file systems already track block pointers), and hence communicating such information requires limited and straightforward modifications to existing infrastructure. More specifically, the modifications required to layers in our approach are *implementation-level*. These modifications are much easier to make compared to the *design-level* modifications required with brand-new abstractions such as Object-based Storage [Mesnier et al. 2003].

(2) By decoupling the *generation* of information at the higher layers from how the information is *used* at the lower layers, we obviate the need for explicit coordination between any two layers to support our abstractions. Our pointer or context information is not generated with any specific layer or functionality in mind.

(3) Our abstractions extend end-to-end across the storage stack, (i.e., from user applications to the storage hardware), hence allowing a wide-range of interesting functionality in the different layers of the storage stack.

We have implemented prototypes of both our abstractions and several case-studies to demonstrate their usefulness, for the Linux kernel 2.6.15. To evaluate disk-level functionality, we built our own software-level disk prototyping framework. Our framework operates as a pseudo device driver that interposes between the file system and the regular disk drivers. One key challenge in this prototyping environment is to ensure there is no performance interference between the host application and the processing at the pseudo driver layer. By careful use of kernel isolation techniques, we separate the CPU and memory usage of the software prototype from the "host" applications, thus providing a very close approximation of an actual hardware prototype with its own processing and memory. We believe that this prototyping environment is valuable more generally for evaluating other kinds of functionality in the storage system. We also plan to release the source code of our framework and the case-studies under GPL.

The key contributions of this article are as follows:

—Formulation of the *pointer* abstraction and the design of the Type-Safe Disk interface that enables easy communication of higher-level pointers to the disk system.

—Design, implementation, and evaluation of two case-studies that demonstrate the security functionality and performance optimizations that type-awareness enables.

—Formulation of the *hierarchical context* abstraction and the Linux implementation of the context propagation infrastructure.

—Design, implementation, and evaluation of two case-studies to demonstrate the power and generality of the context abstraction.

—Implementation of a software-level framework to easily and accurately prototype disk-level functionality. This framework provides an interesting choice between hardware-level prototyping and entirely simulation-based prototyping.

The rest of this article is organized as follows. Section 2 discusses some background information. In Section 3, we present the detailed design, implementation, and evaluation of type-aware storage. In Sections 4 and 5, we describe two case-studies that use type-aware storage. Section 6 presents context-aware I/O. In Sections 7 and 8, we describe two case-studies of context-aware I/O. In Section 9 we discuss related work, and we finally conclude in Section 10.

2. BACKGROUND

In this section, we discuss background information about the modern storage stack, large-scale storage systems, RAID levels, and file systems.

2.1 Modern Storage Stack

In the past file systems communicated directly with disks by using hardware-specific information such as tracks and sectors. The storage stack has evolved significantly since then. Disk hardware information is virtualized through block-based interfaces such as SCSI and ATA. Layers such as RAID [Patterson et al. 1988] or logical volume managers can exist beneath file systems, and they aggregate several independent disks. File systems are completely unaware of whether they are communicating with a single disk system or a RAID array. In today's storage stack, even a network can exist between file systems and the storage hardware [Satran et al. 2004; Sun Microsystems 1989; Callaghan et al. 1995; Shepler et al. 2003], and higher-level user applications are completely oblivous to these intermediate layers.

2.2 Large-Scale Storage Systems

Large-scale storage systems today comprise diverse resources that include high processing power, hundreds of gigabytes of RAM, solid state storage media such as flash, and hundreds or even thousands of disks [EMC Corporation 1999]. Modern storage systems run complex software to provide functionality such as reliability, fault-tolerance, and high performance I/O. One of the challenges in such storage systems is to effectively manage the wide range of resources to provide optimal performance and customizable features. However, despite the advancement in storage hardware, the interface used for communicating with hardware devices is still simple and narrow in most scenarios. For example, the SCSI interface supports just two main primitives, block `read` and `write`, resulting

in the storage system being mostly oblivious to higher-level information. This makes efficient resource management within modern storage systems a difficult problem, as storage systems cannot discriminate between the different kinds of information they store.

Some existing systems try to work around this problem by exporting more information to higher-level software [Denehy et al. 2002; IBM 2007a]. For example, certain enterprise-class storage systems allow higher-level software to choose the RAID level to use for a new volume, during its creation [IBM 2007b]. However, this requires that the file system or higher-level storage software be aware of the characteristics of each volume, which could be totally tied to the internal architecture of the specific storage systems. For example, a storage system could contain several fine-grained RAID levels, and devices such as NVRAM and solid state memory. Storage architectures could also be different across vendors and models, and it may be cumbersome to customize file systems for specific storage systems. Moreover, the abstraction of a volume is in most cases too coarse-grained to express difference in access characteristics across files.

3. TYPE-AWARE STORAGE INFRASTRUCTURE

Type-safety is a well explored concept in the field of programming languages, with proven benefits such as controlled access to memory. We propose to extend the property of type-awareness and type-safety to the disk subsystem, and show that it can significantly improve the security and functionality of the disk subsystem. Specifically, we advocate regulating access to disk blocks to conform to well-defined rules, that are understood and enforced by the disk itself. In building this, we leverage the fact that the semantics of most file systems today can be broadly classified into two categories: raw data blocks, and *pointers* or references that implement logical relationships between data blocks (for example, dentries-inodes and inodes-data blocks). We define a *type-aware* disk as one that can differentiate between these two distinct types of information it stores. Once a disk has this information, it can exploit this knowledge to provide better functionality. We believe that this simple type-awareness could be a significant source of semantic information that can bridge the semantic gap between file systems and storage devices. Although several existing research projects like Object-based Storage Devices (OSD) explore alternatives to bridge this gap, we believe that a *data-pointer* abstraction is the right interface that a disk should provide to file systems. A disk that is type-aware can *enforce* type safety by limiting block accesses to only the legal set of pointers, thus preventing arbitrary block dereferencing. We call such a disk a *type-safe disk* (TSD).

TSDs require a few changes to the current block-based interface. First, like any other type-safe system, allocation and deallocation has to be under the control of the disk system. By performing block allocation and de-allocation, a TSD frees the file system from the need for free-space management. Similar in spirit to type-safe programming languages, a TSD also exploits its pointer awareness to perform automatic garbage collection of unused blocks; blocks which have no pointers pointing to them are reclaimed automatically, thus freeing file systems of the need to track reference counts for blocks in many cases.

In this section we present in more detail, our type-aware storage abstraction, and two case-studies that we built to show the usefulness of our abstraction.

This section is organized as follows. In Section 3.1 we discuss the utility of pointer information at the disk. Section 3.2 discusses the design and implementation of the basic TSD framework. In Section 3.3 we describe file system support for TSDs. In Section 3.4

we present the software-level disk prototyping environment that we built to evaluate the idea of TSDs and all our case-studies. We describe our prototype implementation in Section 3.5. We present the evaluation of our prototype implementation of TSD in Section 3.6.

3.1 Motivation

In this section we present an extended motivation.

Pointers as a proxy for data semantics. The inter-linkage between blocks conveys rich semantic information about the structure imposed on the data by higher layers. Most modern file systems and database systems make extensive use of pointers to organize disk blocks. For example, in a typical file system, directory blocks logically point to inode blocks which in turn point to indirect blocks and regular data blocks. Blocks pointed to by the same pointer block are often semantically related (e.g., they belong to the same file or directory). Pointers also define reachability: if an inode block is corrupt, the file system cannot access any of the data blocks it points to. Thus, pointers convey information about which blocks impact the availability of the file system to various degrees. Database systems are very similar in their usage of pointers. They have B-tree indexes that contain on-disk pointers, and their extent maps track the set of blocks belonging to a table or index.

In addition to being passively aware of pointer relationships, a type-safe disk takes it one step further. It actively enforces invariants on data access based on the pointer knowledge it has. This feature of a TSD enables independent verification of file system operations; more specifically, it can provide an additional perimeter of security and integrity in the case of buggy file systems or a compromised OS. As we show in Section 4, a type-safe disk can limit the damage caused to stored data, even by an attacker with root privileges. We believe this active nature of control and enforcement possible with the pointer abstraction makes it powerful compared to other more passive information-based interfaces.

Pointers thus present a simple but general way of capturing application semantics. By aligning with the core abstraction used by higher-level application designs, a TSD has the potential to enable on-disk functionality that exploits data semantics. In the next subsection, we list a few examples of new functionality (some proposed in previous work in the context of alternative approaches) that TSDs enable.

Applications. There are several possible uses of TSDs.

Selective Data Replication. Since TSDs are capable of differentiating data and pointers, they can identify metadata blocks as those blocks that contain outgoing pointers and replicate them to a higher degree, or distribute them evenly across all the disks. This could provide graceful degradation of availability as provided by D-GRAID [Sivathanu et al. 2004].

Data colocation. Using the knowledge of pointers, a TSD can co-locate blocks along with their reference blocks (blocks that point to them). In general, blocks will be accessed just after their pointer blocks are accessed, and hence there would be better locality during access.

Intelligent Prefetching. TSDs can perform intelligent prefetching of data because of the pointer information. When a pointer block is accessed, a TSD can prefetch the data blocks pointed to by it, and store it in the on-disk buffers for improved read performance.

Disk-level security. TSDs can provide new security properties using the pointer knowledge by enforcing *implicit* capabilities. We discuss this in detail in Section 4.

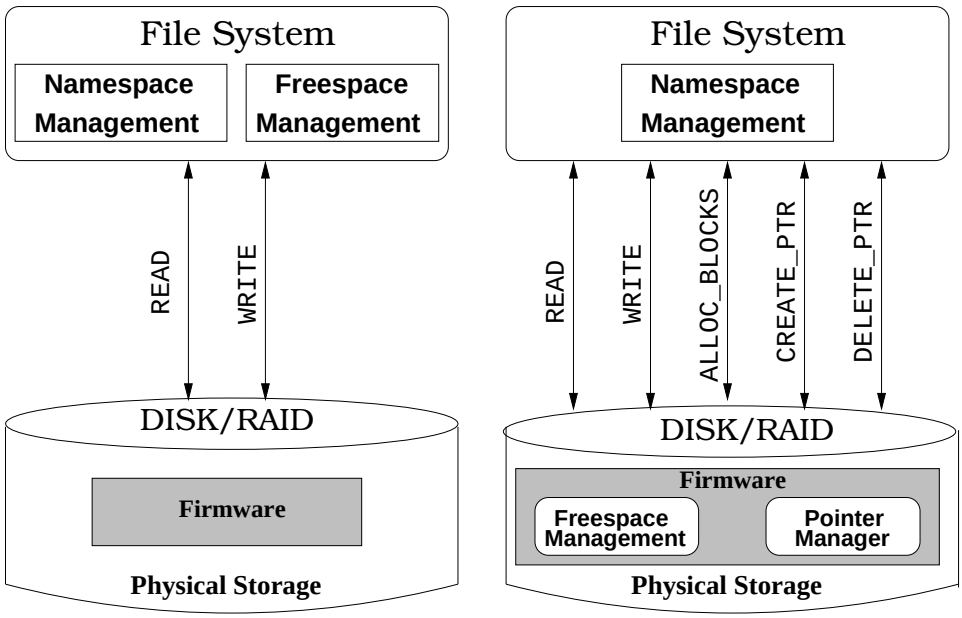
Secure deletion. TSDs can perform automatic secure deletion of deleted blocks by tracking block liveness using pointer knowledge.

3.2 Type-Safety at the Disk Level

Having pointer information inside the disk system enables enforcement of interesting constraints on data access. For example, a TSD allows access to only those blocks that are reachable through some pointer path. TSDs manage block allocations and enforce that every block must be allocated in the context of an existing pointer path, thus preventing allocated blocks from becoming unreachable. More interestingly TSDs enable disk-level enforcement of much richer constraints for data security as described in our case study in section 4.

Enforcing such access constraints based on pointer relationships between blocks is a restricted form of *type-safety*, a well-known concept in the field of programming languages. The type information that a TSD exploits, however, is narrower in scope: TSDs just differentiate between normal data and pointers.

We now detail the TSD interface, its operation, and our prototype implementation. Figure 1 shows the architectural differences between normal disks and a TSD.



(a) Traditional Disk (b) Type−safe Disk

Fig. 1. *Comparison of traditional disks vs. type-safe disks*

Disk API. A type-safe disk exports the following primitives, in addition to the basic block-based API:

—SET_BLOCKSIZE($Size$): Sets the file system block size in bytes.

—ALLOC_BLOCKS($Ref, Hint, Count$): Allocates $Count$ number of new file system blocks from the disk-maintained free block list, and creates pointers to the allocated blocks, from block Ref. Allocated blocks need not be contiguous. Ref must be a valid block number that was previously allocated. $Hint$ is the block number closest to which the new blocks should be allocated. $Hint$ can be NULL, which means the disk can choose the new block totally at its own discretion. Returns an array of addresses of the newly allocated blocks, or NULL if there are not enough free blocks on the device.

—ALLOC_CONTIG_BLOCKS($Ref, Hint, Count$): Follows the same semantics as ALLOC_BLOCKS, except that it allocates $Count$ number of contiguous blocks if available.

—CREATE_PTR($Src, Dest$): Creates a pointer from block Src to block $Dest$. Both Src and $Dest$ must be previously allocated. Returns success or failure.

—DELETE_PTR($Src, Dest$): Deletes a pointer from block Src that points to block $Dest$. Semantics similar to CREATE_PTR.

—GET_FREE: Returns the number of free blocks left.

Managing Block Pointers. A TSD needs to maintain internal data-structures to keep track of all pointers between blocks. It maintains a pointer tracking table called PTABLE that stores the set of all pointers. The PTABLE is indexed by the source block number and each table entry contains the list of destination block numbers. A new PTABLE entry is added every time a pointer is created. Based on pointer information, TSD disk blocks are classified into three kinds: (a) *Reference blocks*: blocks with both incoming and outgoing pointers (such as inode blocks). (b) *Data blocks*: blocks without any outgoing pointers but just incoming pointers. (c) *Root blocks*: a pre-determined set of blocks that contain just outgoing pointers but not incoming pointers. Root blocks are never allocated or freed, and they are statically determined by the disk. Root blocks are used for storing boot information or the primary metadata block of file systems (e.g., the Ext2 super block).

Free-Space Management. To perform free-space management at the disk level, we track live and free blocks. A TSD internally maintains an allocation bitmap, ALLOC-BITMAP, containing one bit for every logical unit of data maintained by the higher level software (e.g., a file system block). The size of a logical unit is set by the upper-level software through the SET_BLOCKSIZE disk primitive. When a new block need to be allocated, the TSD can choose a free block closest to the hint block number passed by the caller. Since the TSD can exploit the low level knowledge it has, it chooses a block number which requires the least access time from the hint block.

TSDs use the knowledge of block liveness (a block is defined to be dead if it has no incoming pointers) to perform garbage collection. Unlike traditional garbage collection systems in programming languages, garbage collection in TSD happens *synchronously* during a particular DELETE_PTR call which deletes the last incoming pointer to a block. A TSD maintains a reference count table, RTABLE, to speed up garbage collection. The reference count of a block gets incremented every time a new incoming pointer is created and is decremented during pointer deletions. When the reference count of a block drops to zero during a DELETE_PTR call, the block is marked free immediately. A TSD performs garbage collection one block at a time as opposed to performing cascading deletes.

Garbage collection of reference blocks with outgoing pointers is prevented by disallowing deletion of the last pointer to a reference block before all outgoing pointers in it are deleted.

Consistency. As TSDs maintain separate pointer information, TSD pointers could become inconsistent with the file system pointers during system crashes. Therefore, upon a system crash, the consistency mechanism of the file system is triggered which checks file system pointers against TSD pointers and first fixes any inconsistencies between both. It then performs a regular scan of the file system to fix file system inconsistencies and update the TSD pointers appropriately. For example, if the consistency mechanism creates a new inode pointer to fix an inconsistency, it also calls the CREATE_PTR primitive to update the TSD internal pointers. Alternatively, we can obviate the need for consistency mechanisms by just modifying file systems to use TSD pointers instead of maintaining their own copy in their meta-data. However, this involves wide-scale modifications to the file system.

File system integrity checkers such as `fsck` for TSDs have to run in a privileged mode so that they can perform a scan of the disk without being subjected to the constraints enforced by TSDs. This privileged mode can use a special administrative interface that overrides TSD constraints and provides direct access to the TSD pointer management data-structures.

Block corruption. When a block containing TSD-maintained pointer data-structures gets corrupted the pointer information has to be recovered, as the data blocks pertaining to the pointers could still be reachable through the file system meta-data. Block corruption can be detected using well-known methods such as checksumming. Upon detection, the TSD notifies the file system, which recreates the lost pointers from its meta-data.

3.3 File System Support

We now describe how a file system needs to be modified to use a TSD. We first describe the general modifications required to make any file system work with a TSD. Next, we describe our modifications to two file systems, Linux Ext2 and VFAT, to use our framework.

Since TSDs perform free-space management at the disk-level, file systems using TSDs are freed from the complexity of allocation algorithms, and tracking free block bitmaps and other related meta-data. However, file systems now need to call the disk API to perform allocations, pointer management, and getting the free blocks count. The following are the general modifications required to existing file systems to support type-safe disks:

(1) The `mkfs` program should set the file system block size using the SET_BLOCKSIZE primitive, and store the primary meta-data block of the file system (e.g., the Ext2 super block) in one of the TSD root blocks. Note that the TSD root blocks are a designated set of well-known blocks known to the file system.

(2) The free-space management sub-system should be eliminated from the file system, and TSD API should be used for block allocations. The file system routine that estimates free-space, should call the GET_FREE disk API, instead of consulting its own allocation structures.

(3) Whenever file systems add new pointers to their meta-data, CREATE_PTR disk primitive should be called to create a TSD pointer. Similarly, the DELETE_PTR primitive has to be called when pointers are removed from the file system.

In the next two sub-sections we describe the modifications that we made to the Ext2 and the VFAT file systems under Linux, to support type-safe disks.

Ext2TSD. We modified the Linux Ext2 file system to support type-safe disks; we call the modified file system *Ext2TSD*. The Ext2 file system groups together a fixed number of sequential blocks into a block group and the file system is managed as a series of block groups. This is done to keep related blocks together. Each block group contains a copy of the super block, inode and block allocation data-structures, and the inode blocks. The inode table is a contiguous array of blocks in the block group that contain on-disk inodes.

To modify Ext2 to support TSDs, we removed the notion of block groups from Ext2. Since allocations and de-allocations are done by using the disk API, the file system need not group blocks based on their order. However, to perform easy inode allocation in tune with Ext2, we maintain inode groups which we call ISEGMENTS. Each isegment contains a segment descriptor that has an inode bitmap to track the number of free inodes in that isegment. The inode allocation algorithm of Ext2TSD is same as that of Ext2. The `mkfs` user program of Ext2TSD writes the super block, and allocates the inode segment descriptor blocks, and inode tables using the allocation API of the disk. It also creates pointers from the super block to all blocks containing isegment descriptors and inode tables.

The organization of file data in Ext2TSD follows the same structure as Ext2. When a new file data or indirect block is allocated, Ext2TSD calls ALLOC_BLOCKS with the corresponding inode block or the indirect block as the reference block. While truncating a file, Ext2TSD just deletes the pointers in the indirect block branches in the right order such that all outgoing pointers from the parent block to its child blocks are deleted before deleting the incoming pointer to the parent block. Thus blocks belonging to truncated or deleted files are automatically reclaimed by the disk.

In the Ext2 file system, each directory entry contains the inode number for the corresponding file or directory. This is a logical pointer relationship between the directory block and the inode block. In our implementation of Ext2TSD, we create physical pointers between a directory block and the inode blocks corresponding to the inode numbers contained in every directory entry in the directory block. Modifying the Ext2 file system to support TSD was relatively simple. It took 8 days for us to build Ext2TSD starting from a vanilla Ext2 file system. We removed 538 lines of code from Ext2 which are mostly the code required for block allocation and bitmap management. We added 90 lines of new kernel code and modified 836 lines of existing code.

3.4 A Software-Level Disk Prototyping Framework

In this section, we describe our generic disk functionality prototyping framework, DPROTO, that we built for the Linux kernel 2.6.15.

We developed DPROTO as a pseudo-device driver that stacks on top of one or more lower-level disk or software RAID drivers, in a single machine. One of the main challanges in developing DPROTO is isolating the resources consumed by components that are supposed to go inside the disk firmware if it were a real implementation. For example, if the functionality being prototyped is a disk-level data compression technique, the part of DPROTO that performs compression has to consume resources that are completely isolated from that used by applications and file systems, which is difficult in a single machine setup.

While developing DPROTO we aimed at isolating key resources, CPU and memory, between disk-level functionality and higher-level applications. For CPU isolation, we use a multiprocessor setup and ensure that disk-level functionality always gets executed in a dedicated processor. For memory isolation, we implemented an isolated preallocated memory

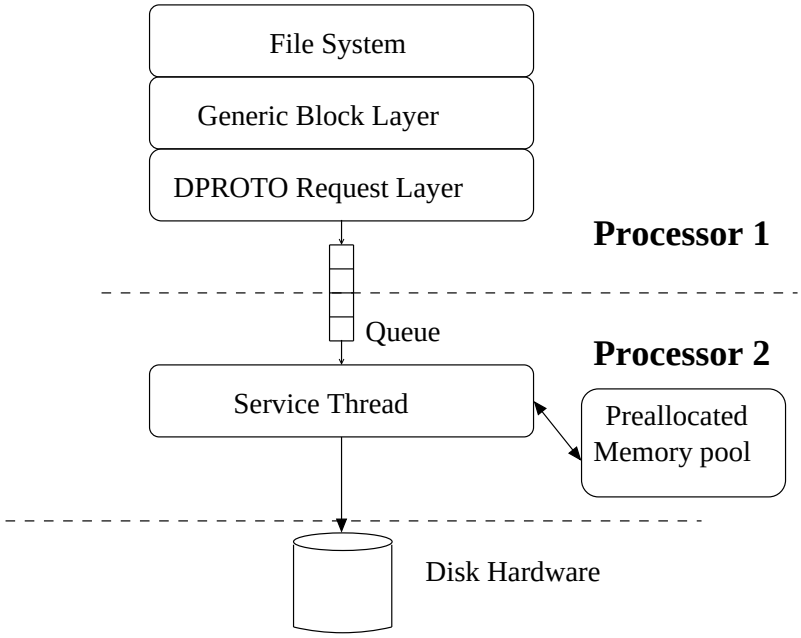


Fig. 2. *DPROTO Architecture*

pool and ensured that disk functionality never accesses memory beyond the preallocated range.

Figure 2 shows the architecture of DPROTO. We implemented the pseudo-device driver as two layers: the upper layer running in the context of the file system and the lower layer running as a separate thread bound to an isolated CPU. Disk I/O requests generated from the file system reach the upper layer of DPROTO, which adds the request to a shared queue. The lower layer services requests from the queue and eventually passes it down to physical storage. Any disk-level functionality such as compression would be handled by the lower-level service thread and hence runs in an isolated CPU. All memory allocations done by both layers of DPROTO use the preallocated memory pool. Therefore, DPROTO requires specifying the total memory requirement for a given functionality before hand.

To test the performance of a disk-level functionality prototyped using DPROTO, the comparison reference can be run with one processor disabled and with the appropriate size of memory preallocated. For example, if a compression disk system is compared to a regular disk system for a particular workload, the regular disk run of the workload has to be done with one processor disabled and the preallocated memory equal to the memory requirement of the compression disk. With this procedure, the comparison becomes fair and closely represents the results of a real implementation.

Our implementation of DPROTO had 5,790 lines of new kernel code and 350 lines of user-level code.

DPROTO Overheads. We evaluated the performance of DPROTO framework as a null layer that stacks on top of a regular disk. We ran Postmark for two different configurations on an Ext2 file system mounted on the null DPROTO layer, and compared it with Postmark

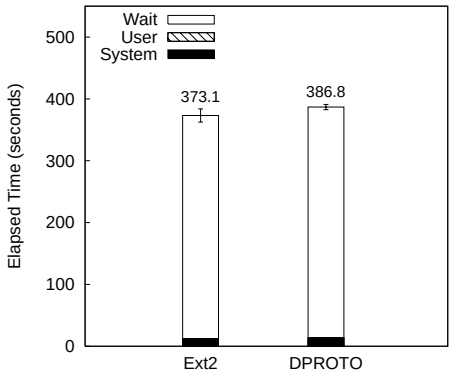


Fig. 3. Postmark results for DPROTO vs. regular disk

run on a regular disk. Figure 3 shows the overheads of DPROTO compared to a regular disk. The overall elapsed time overhead of DPROTO was 3.6% compared to regular disks. This is contributed mostly by increased in wait time, caused because of an additional level of indirection in the DPROTO request service queue.

3.5 TSD Implementation

We implemented a prototype TSD using our DPROTO software-level disk prototyping framework, in the Linux kernel 2.6.15. It contains 3,108 lines of kernel code. The TSD layer receives all block requests, and redirects the common read and write requests to the lower level device driver. The additional primitives required for operations such as block allocation and pointer management are implemented as driver `ioctl`s.

We implemented PTABLE and RTABLE as in-memory hash tables which gets written out to disk at regular intervals of time through an asynchronous commit thread. In implementing the RTABLE, we add an optimization to reduce the number of entries maintained in the hash table. We add only those blocks whose reference count is greater than one. A block which is allocated and which does not have an entry in the RTABLE is deemed to have a reference count of one and an unallocated block (as indicated by the ALLOC_BITMAP) is deemed to have a reference count of zero. This significantly reduces the size of our RTABLE, because most disk blocks have reference counts of zero or one (e.g., all data blocks have reference counts zero or one).

Memory usage. In our prototype implementation we maintained all TSD data-structures in memory. The space overheads associated with TSD pointer tracking and free-space management is directly related to the number of file system blocks on disk. We found that the TSD pointer meta-data per file system block will be close to 20 bytes (with an average of one incoming pointer per block). Assuming a file system block size of 4KB, the total space overheads for TSDs totals upto 0.5% of the disk size. In a real firmware-level implementation of TSDs, the entire meta-data need not be maintained in memory. At any given time, the working-set of TSD pointers is limited to the directories and files being accessed. Hence, we believe that it would be sufficient if a fraction of the TSD meta-data (about 10%) is cached in memory, and the rest of the meta-data can be stored on secondary storage.

3.6 Evaluation

We evaluated the performance of our prototype TSD framework in the context of Ext2TSD. We ran general-purpose workloads and also micro-benchmarks on our prototype and compared them with unmodified Ext2 file system on a regular disk. This section is organized as follows: first we talk about our test platform, configurations, and procedures. Next, we analyse the performance of the TSD framework with the Ext2TSD file system.

Test Infrastructure. We conducted all tests on a 2.8GHz Xeon with 1GB RAM, and a 250GB, LSILogic SCSI disk. We used Fedora Core 6, running a vanilla Linux 2.6.15 kernel. To ensure a cold cache, we unmounted all involved file systems between each test. We ran all tests at least five times and computed 95% confidence intervals for the mean elapsed, system, user, and wait times using the Student-t distribution. In each case, the half-widths of the intervals were less than 5% of the mean. Wait time is the elapsed time less CPU time used and consists mostly of I/O, but process scheduling can also affect it. We recorded disk statistics from `/proc/diskstats` for our test disk. We analysed the following detailed disk-usage statistics while interpreting the results: the number of read I/O requests (`rio`), number of write I/O requests (`wio`), number of sectors read (`rsect`), number of sectors written (`wsect`), number of read requests merged (`rmerge`), number of write requests merged (`wmerge`), total time taken for read requests (`ruse`), and the total time taken for write requests (`wuse`).

Benchmarks and Configurations.

Postmark. We used Postmark v1.5 to generate an I/O-intensive workload. Postmark stresses the file system by performing a series of operations such as directory lookups, creations, and deletions on small files [Katcher 1997]. Postmark is typically configured by specifying a number of initial files, and a fixed number of *transactions*. Postmark then creates the initial pool of files, performs the fixed number of transactions, and removes any left over files.

Kernel Compile. To simulate a relatively CPU-intensive user workload, we compiled the Linux kernel source code. We used a vanilla Linux 2.6.15 kernel, and analyzed the overheads of Ext2TSD, for the `make oldconfig` and `make` operations combined.

Sprite LFS Benchmarks. To isolate the overheads of individual file system operations, we ran the entire suite of Sprite LFS benchmarks [Rosenblum 1992]. The Sprite LFS benchmarks contains two sets of workloads, for meta-data and data operations. The first set deals with small files and tests, file creation, read, and file deletion. The second set operates on large files and performs sequential and random reads and writes.

Postmark Results. We ran the Postmark benchmark on three setups: (1) regular Ext2 over a regular disk, (2) regular Ext2 on DPROTO, and (3) Ext2TSD over our implementation of TSD. We configured Postmark with two different configurations. In the first configuration, we used 10,000 files with sizes ranging from 100KB to 200KB, and 10,000 transactions. Figure 4(a) shows the overheads of DPROTO and the TSD infrastructure for this configuration. As evident from the figure, Ext2 over our prototyping infrastructure DPROTO had negligible overheads compared to Ext2 over a regular disk. However Ext2TSD ran 7% faster than regular Ext2 inspite of a 1.3 times increase in system time. The increase in system time is because of the device `ioctls` that Ext2TSD calls for the

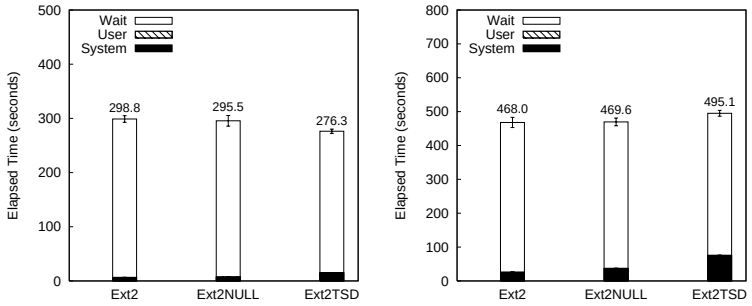


Fig. 4. *Postmark Results: (a) 10,000 files, sizes 100KB to 200KB, 10,000 transactions. Ext2NULL indicates the results for regular Ext2 over a NULL pseudo-device driver. (b) 1,000 files, sizes 1MB to 3MB, 5,000 transactions.*

pointer operations. From the kernel disk I/O statistics, we found that the 10% decrease in wait time for Ext2TSD compared to regular Ext2 is caused by more requests getting merged at the device driver layer. This is because, block allocation is performed by TSDs in the case of Ext2TSD, and there was better spatial locality compared to regular Ext2.

Figure 4(b) shows the results for a different configuration of Postmark. For this we used 1,000 files with sizes ranging from 1MB to 3MB, and performed 5,000 transactions. In this configuration, Ext2TSD had an elapsed time overhead of 5.7% compared to regular Ext2. The system time overhead was 1.9 times and wait time was 5% lesser than regular Ext2. This shows that for larger files, the savings in I/O time because of better spatial locality is lesser compared to smaller files.

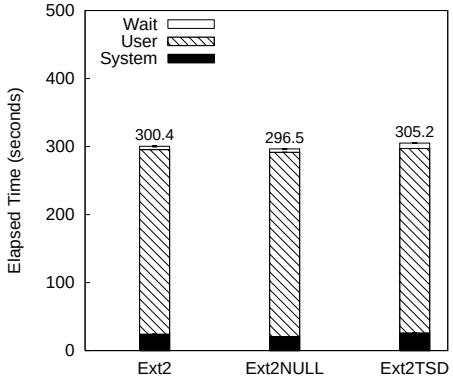


Fig. 5. *Kernel Compile Results. Ext2NULL indicates the results for regular Ext2 over a NULL pseudo-device driver.*

Kernel Compile Results. Results for the kernel compilation benchmark is shown in Figure 5. Ext2TSD had a small elapsed time overhead of 1.5% compared to regular Ext2. This was caused by a 7% increase in system time and 60% increase in wait time. The system time increase in this case is smaller compared to the Postmark results because kernel compile is a predominantly CPU-intensive workload and hence has much lesser number of

pointer operations. The wait time increase is because the main compilation thread waits for the DPROTO disk thread to complete pointer operations. The wait time increase is more pronounced here because the time interval between I/O is larger than that of Postmark.

Sprite LFS Benchmark Results. We ran the entire suite of Sprite LFS benchmarks on Ext2 over a regular disk, and Ext2TSD over our prototype TSD.

Meta-data benchmarks. To generate a small file creation workload, we created 1,000,000 files, with size 4KB each, in 1,000 sub-directories. For reads, we remounted the file system and read all the 1,000,000 files we created. For deletes, we unlinked all files.

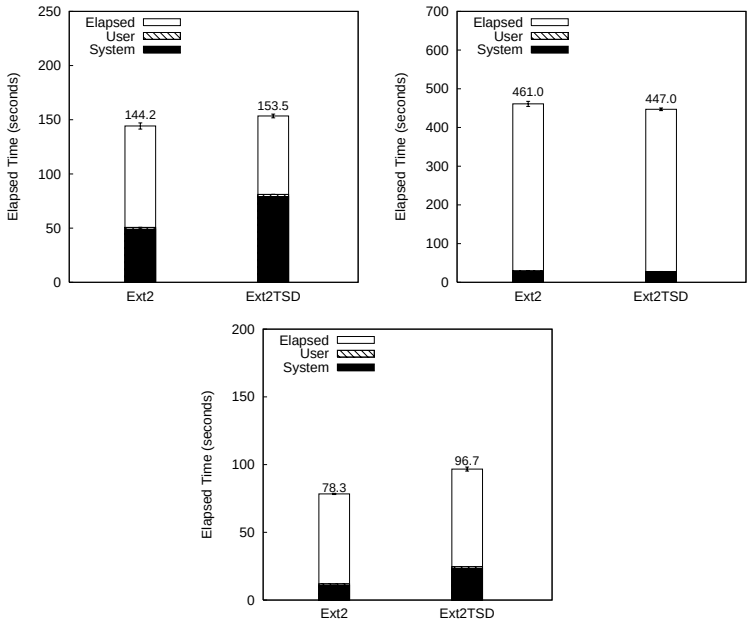


Fig. 6. *Sprite LFS benchmarks: (a) Create results. (b) Read results. (c) Delete results.*

Figure 6(a) shows the overheads of Ext2TSD. Ext2TSD had an elapsed time overhead of 6.4% compared to regular Ext2. This is because of a 61% increase in system time. The system time increase is because of the pointer operations as this workload consists of intensive meta-data write operations. The wait time mainly caused by I/O, reduced by 22% because the TSD allocation policy is favorable for small files.

Figure 6(b) shows the results of the read workload. Ext2TSD performed 3% better than regular Ext2. The system time reduced by 5% because of two reasons. First, there are no pointer operations in a read workload. Second, the isolation technique in DPROTO offloads part of call stack of I/O operations such as lower level SCSI driver calls, to the DPROTO disk thread.

The delete workload results shown in Figure 6(c) shows that the elapsed time overhead of Ext2TSD is 23% compared to regular Ext2. This is caused because of a 2.1 times increase in system time. This increase is because of a large number of pointer deletion operations happening within a short period of time.

Overall, even under extremely meta-data intensive workloads, the elapsed time overheads are moderate. In most common environments such meta-data intensive workloads are unlikely.

Data benchmarks. For generating Sprite LFS data benchmark workloads, we used a large file of size 4GB. For random workloads we performed 10,000 random 4K reads or writes. To eliminate cache effects, we generated a duplicate free list of random page numbers. For sequential workloads, we performed 1,000,000 sequential 4K reads on the file.

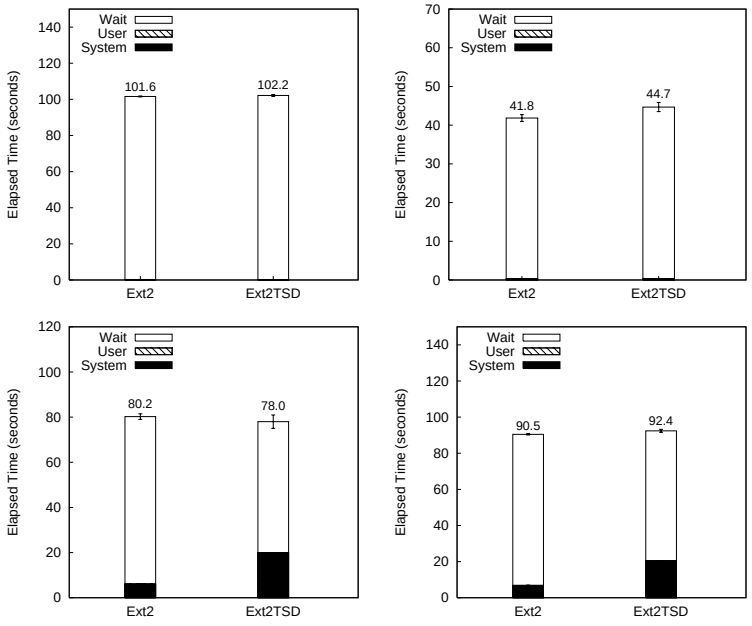


Fig. 7. *Sprite LFS Data benchmark results: (a) Random read. (b) Random write. (c) Sequential Read. (d) Sequential Write.*

Figure 7(a) shows the results for the random read workload. Ext2TSD had no visible overheads for this. As this is a read workload, it generated no pointer operations. For random write, as shown in Figure 7(b), Ext2TSD had an elapsed time overhead of 6.7%. This is mainly caused by a 6.7% increase in wait time. The wait time increase is because the main benchmark thread had to wait for the DPROTO disk thread to service pointer operations.

Figure 7(c) shows the results for sequential read. The difference in elapsed time between Ext2TSD and regular Ext2 was negligible. However, the system overhead in Ext2TSD was 2.3 times. This was offset by a 21% reduction in wait time. As this is a sequential workload, a very large number I/O operations were executed within a short time interval. This resulted in making CPU overheads more visible. The CPU overheads were due to lock contention for the request queue shared by the main benchmark thread and the DPROTO disk thread. Our implementation uses a `spin_lock` for this, and hence it shows up as

system time. The wait time decrease is because of better spatial locality in the case of Ext2TSD.

Figure 7(d) shows the results for sequential writes. The overheads of Ext2TSD were similar to sequential reads, as our sequential write workload performed overwrites of existing file data, resulting no additional pointer operations.

In summary, our evaluation shows that the overheads associated with our TSD disk infrastructure and the Ext2TSD file system is quite minimal (about 2%) for normal user workloads. This is shown by the results of our kernel compilation benchmark. For more I/O-intensive workloads such as Postmark and Sprite meta-data benchmarks, Ext2TSD shows overheads as high as 23%. We used such benchmarks to show the worst case overheads of TSDs. However, such I/O-intensive workloads are uncommon in real scenarios. Most of the system-time overheads were caused by pointer operations issued by the file system. This could be reduced by aggregating the operations and sending it to the disk system in batches. While the allocation primitive has to be synchronous, pointer creation and deletion can be made asynchronous.

4. CASE STUDY: ACCESS

We describe how type-safety can enable a disk to provide better security properties than existing storage systems. We designed and implemented a secure storage system called ACCESS (***A C**apability **C**onscious **E**xtended **S**torage **S**ystem*) using the TSD framework; we then built a file system on top, called Ext2ACCESS. We first motivate the need for enforcing disk-level capabilities, then present a detailed design of ACCESS. Finally, we describe our prototype implementation of ACCESS and the implementation of Ext2ACCESS, a file system that supports ACCESS.

Protecting data confidentiality and integrity during intrusions is crucial: attackers should not be able to read or write on-disk data even if they gain root privileges. One solution is to use encryption [Blaze 1993; Wright et al. 2003]; this ensures that intruders cannot decipher the data they steal. However, encryption does not protect the data from being overwritten or destroyed. An alternative is to use explicit disk-level *capabilities* to control access to data [Aguilera et al. 2003; Gibson et al. 1998]. By enforcing capabilities independently, a disk enables an additional perimeter of security even if the OS is compromised. Others explored using disk-level versioning that never overwrites blocks, thus enabling the recovery of pre-attack data [Strunk et al. 2000].

ACCESS is a type-safe disk that uses pointer information to enforce *implicit path-based* capabilities, obviating the need to maintain explicit capabilities for all blocks, yet providing similar guarantees.

ACCESS has five design goals. (1) Provide an infrastructure to limit the scope of confidentiality breaches on data stored on local disks even when the attacker has root privileges or the OS and file systems are compromised. (2) The infrastructure should also enable protection of stored data against damage even in the event of a network intruder gaining access to the raw disk interface. (3) Support efficient and easy revocation of authentication keys, which should not require costly re-encryptions upon revocation. (4) Enable applications to use the infrastructure to build strong and easy-to-use security features. (5) Support data recovery through administrative interfaces even when authentication tokens are lost.

4.1 Design

The primitive unit of storage in today's commodity disks is a fixed-size disk block. Authenticating every block access using a capability is too costly in terms of performance and usability. Therefore, there needs to be some criteria by which blocks are grouped and authenticated together. Since TSDs can differentiate between normal data and pointers, they can perform logical grouping of blocks based on the reference blocks pointing to them. For example, in Ext2 all data blocks pointed to by the same indirect block belong to the same file.

ACCESS provides the following guarantee: a block x cannot be accessed unless a valid reference block y that points to this block x is accessed. This guarantee implies that protecting access to data simply translates to protecting access to the reference blocks. Such grouping is also consistent with the fact that users often arrange files of related importance into individual folders. Therefore, in ACCESS, a single capability would be sufficient to protect a logical working set of user files. Reducing the number of capabilities required is not only more efficient, but also more convenient for users.

In ACCESS, blocks can have two capability strings: a `read` and a `write` capability (we call these *explicit capabilities*). Blocks with associated explicit capabilities, which we call *protected* blocks, can be read or written only by providing the appropriate capability. By performing an operation on a block Ref using a valid capability, the user gets an *implicit capability* to perform the same operation on all blocks pointed to by Ref, which are not directly protected (capability inheritance). If a particular reference block i points to another block j with associated explicit capabilities, then the implicit capability of i is not sufficient to access j; the explicit capability of j is needed to perform operations on it.

As all data and reference blocks are accessed using valid pointers stored on disk, root blocks are used to bootstrap the operations. In ACCESS, there are three kinds of access modes: (1) All protected blocks are accessed by providing the appropriate capability for the operation. (2) Blocks which are not protected can inherit their capability from an authenticated parent block. (3) Root blocks can be accessed without any reference block by providing the appropriate capability, if they are protected.

ACCESS Meta-Data. ACCESS maintains a table named KTABLE indexed by the block number, to store the blocks' `read` and `write` capabilities. During every block access it checks if the block has a KTABLE entry. If there is a KTABLE entry, the capability provided by the user is authenticated against the stored capability before performing the operation. ACCESS tracks the list of all reference blocks that are accessed successfully in a given period of time, and uses it to authenticate accesses to the blocks that do not have associated capabilities.

ACCESS also maintains a temporal access table called LTABLE which is indexed by the reference block number. The LTABLE has entries for all reference blocks whose associated implicit capabilities have not timed out. The timed out entries in the LTABLE are periodically purged.

Preventing Replay Attacks. In ACCESS, data needs to be protected even in situations where the OS is compromised. Passing clear-text capabilities through the OS interface could lead to replay attacks by a silent intruder who eavesdrops capabilities. To protect against this, ACCESS associates a sequence number with capability tokens. To read a protected block, the user has to provide a HMAC checksum of the capability (C_u) con-

catenated with a sequence number (S_u) ($H_u = HMAC(C_u + S_u, C_u)$). This can be generated using an external key card or a hand-held device that shares sequence numbers with the ACCESS disk system. Each user has one of these external devices, and ACCESS tracks sequence numbers for each user's external device. Upon receiving H_u for a block, ACCESS retrieves the capability token for that block from the KTABLE and computes $H_A = HMAC(C_A + S_A, C_A)$, where C_A and S_A are the capability and sequence number for the block, and are maintained by ACCESS. If H_u and H_A do not match, ACCESS denies access. Skews in sequence numbers are handled by allowing a window of valid sequence numbers at any given time.

ACCESS Operation. During every reference block access, an optional timeout interval ($Interval$) can be provided, during which the implicit capabilities associated with that reference block will be active. Whenever a reference block Ref is accessed successfully, an LTABLE entry is added for it. This entry stays until $Interval$ expires. It is during this period of time, that we call the *temporal window*, all child blocks of Ref which are not protected inherit the implicit capability of accessing Ref. Once the timeout interval expires, all further accesses to the child blocks are denied. This condition should be captured by the upper level software, which should prompt the user for the capability token, and then call the disk primitive to renew the timeout interval for Ref. The value of $Interval$ can be set based on the security and convenience requirements. Long-running applications that are not interactive in nature should choose larger timeout intervals.

At any instant of time when the OS is compromised, the subset of blocks whose temporal window is active will be vulnerable to attack. This subset would be a small fraction of the entire disk data. The amount of data vulnerable during OS compromises can be reduced by choosing short timeout intervals. One can also force the timeout of the temporal window using the FORCE_TIMEOUT disk primitive described below.

To read a data block in ACCESS, the base pointer should be read first from one of the root blocks, by presenting the appropriate capability. If the access of the root block is successful, ACCESS will add an entry for the root block in the LTABLE. Once this is done, all blocks pointed to by the root block that do not have associated capabilities can be accessed until the LTABLE entry times out. In the context of a file system, the initial root block read would be its super block, and this occurs during `mount`. The temporal locality of the initial super block access is used as an implicit capability for accessing subsequent blocks. Whenever an implicit capability for a block needs to be verified, the disk checks if the reference block passed by the upper level software has an LTABLE entry for it. If an entry does not exist, ACCESS denies access to the block. If the reference block has an LTABLE entry, ACCESS looks up the PTABLE to find if the reference block indeed has a pointer to the block whose implicit capability needs to be verified. The reference block passed by the upper level software is only used for optimizing performance during the temporal lookup.

For blocks with associated capabilities, the appropriate capability string must be provided. Each reference block can have its own read and write capabilities depending on the owner of that reference block. For example, an indirect block of a particular user's file will have that user's capabilities, and cannot be read by anyone other than that person.

ACCESS API. To design the ACCESS API, we extended the TSD API (Section 3.2) with capabilities, and added new primitives for managing capabilities and timeouts. Note

that some of the primitives described below let the file system specify the reference block through which the implicit capability chain is established. However, as we describe later, this is only used as a hint by the disk system for performance reasons; ACCESS maintains its own structures that validate whether the specified reference block was indeed accessed, and it has a pointer to the actual block being accessed. In this section when we refer to read or write *capabilities*, we mean the HMAC of the corresponding capabilities and a sequence number.

(1) SET_CAPLEN($Length$): Sets the length of capability tokens. This setting is global.

(2) ALLOC_BLOCKS($Ref, Ref_r or C_w, Count$): Operates similar to the TSD ALLOC_BLOCKS primitive with the following two changes. (1) If Ref is protected the call takes the write capability of Ref, C_w; (2) otherwise, the call takes the reference block Ref_r of Ref, to verify that the caller has write access to Ref.

(3) ALLOC_CONTIG_BLOCKS($Ref, Ref_r or C_w, Count$): Same as the ALLOC_BLOCKS primitive, but allocates contiguous blocks.

(4) READ($Bno, Ref or C_{rw}, Timeout$): Reads the block represented by Bno. Ref is the reference block that has a pointer to Bno. C_{rw} is either the read or the write capability of block Bno. The second argument of this primitive must be Ref if Bno is not protected for read, and must be C_{rw} if Bno is protected. $Timeout$ is the timeout interval.

(5) WRITE($Bno, Ref or C_w, timeout$): Writes the block represented by Bno. C_w is the write capability of Bno. Other semantics are similar to READ.

(6) CREATE_PTR($Src, Dest, Ref_s or C_{sw}, C_{dw} or Ref_{dw}$): Creates a pointer from block Src to block $Dest$. If Src or $Dest$ are protected, their capabilities have to be provided. For blocks which are not protected, the caller must provide valid reference blocks which point to Src and $Dest$. Note that although the pointer is created only from the source block, we need the write capability for the destination block as well; without this requirement, one can create a pointer to any arbitrary block and gain implicit write capabilities on that block.

(7) DELETE_PTR($Src, Dest, Ref_s or C_{sw}$): Deletes a pointer from block Src to block $Dest$. Write credentials for Src has to be provided.

(8) KEY_CONTROL($Bno, C_{ow}, C_{nr}, C_{nw}, Ref$): This sets, unsets, or changes the read and write capabilities associated with the block Bno. C_{ow} is the old write capability of Bno. C_{nr} and C_{nw} are the new read and write capabilities respectively. A reference block Ref that has a pointer to Bno needs to be passed only while setting the write key for a block that did not have a write capability before. For all other operations, like unsetting keys or changing keys, Ref need not be specified because C_{ow} can be used for authentication.

(9) RENEW_CAPABILITY($Ref, C_{rw}, Interval$): Renews the capability for a given reference block. C_{rw} is the read or write key associated with Ref. $Interval$ is the timeout interval for the renewal.

(10) FORCE_TIMEOUT(Ref): Times out the implicit capabilities associated with reference block Ref.

(11) SET_BLOCKSIZE and GET_FREE TSD primitives (Section 3.2) can be called through a secure administrative interface.

4.2 ACCESS Prototype

We extended our TSD prototype to implement ACCESS. We implemented additional hash tables for storing the KTABLE and LTABLE required for tracking capabilities and temporal access locality respectively. All in-memory hash tables were periodically committed to disk through an asynchronous commit thread. The allocation and pointer management `ioctl`s in TSD were modified to take capabilities or reference blocks as additional arguments. We implemented the KEY_CONTROL primitive as a new `ioctl` in our pseudo-device driver.

To authenticate the `read` and `write` operations, we implemented a new `ioctl`, KEY_INPUT. We did this to simplify our implementation and not modify the generic block driver. The KEY_INPUT `ioctl` takes the block number and the capabilities (or reference blocks) as arguments. The upper level software should call this `ioctl` before every read or write operation to authenticate the access. Internally, the disk validates the credentials provided during the `ioctl` and stores the success or failure state of the authentication. When a read or write request is received, ACCESS checks the state of the previous KEY_INPUT for the particular block to allow or disallow access. Once access is allowed for an operation, the success state is reset. When a valid KEY_INPUT is not followed by a subsequent read or write for the block (e.g., due to software bugs), we time out the success state after a certain time interval. This method of using an `ioctl` for sending the credentials greatly simplified our prototype implementation, as we did not have to modify the generic block driver interfaces to send additional arguments during the read and write operations.

4.3 The Ext2ACCESS File System

We modified the Ext2TSD file system described in Section 3.3 to support ACCESS; we call the new file system *Ext2ACCESS*. To demonstrate a usage model of ACCESS disks, we protected only the inode blocks of Ext2ACCESS with read and write capabilities. All other data blocks and indirect blocks had implicit capabilities inherited from their inode blocks. This way users can have a single read or write capability for accessing a whole file. An alternative approach may be to protect only directory inode blocks. ACCESS provides an infrastructure for implementing security at different levels, which upper level software can use as needed.

4.4 Evaluation

We evaluated the performance of ACCESS using our Ext2ACCESS file system. We compared Ext2ACCESS with a regular Ext2 file system mounted on a regular disk. The hardware setup we used was same as that for evaluating the TSD infrastructure, described in Section 3.6. We ran two different workloads: Postmark, kernel compilation as described below.

Postmark Results. Figure 8 shows the results for Postmark. For this benchmark, we configured Postmark with 10,000 files of sizes ranging from 100KB to 200KB, and 10,000 transactions. Ext2ACCESS performed 19% better than regular Ext2, mainly because of a 24% decrease in I/O time. The difference in I/O time in this case is more than that of Ext2TSD vs. regular Ext2 discussed in Section 3.6 because ACCESS pre-allocates more memory than regular TSD for its data-structures. This results in reduced cache size making the impact of spatial locality more pronounced. The system time for Ext2ACCESS was 3 times more than that of regular Ext2 mainly because of pointer and key management

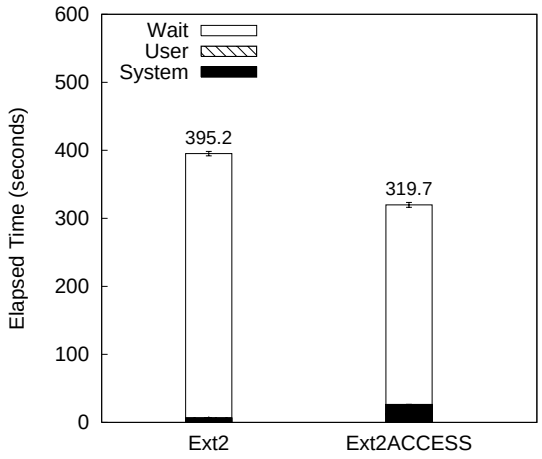


Fig. 8. *Postmark Results for ACCESS*

`ioctl`s issued by Ext2ACCESS.

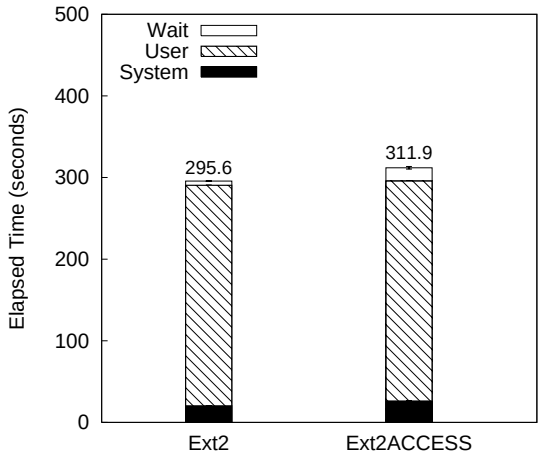


Fig. 9. *Kernel Compile Results for ACCESS*

Kernel Compile Results. Figure 9 shows the kernel compilation results for Ext2ACCESS. As evident from the figure, the overall elapsed time overhead of Ext2ACCESS was 5% compared to regular Ext2. This is caused by a 29% increase in system time and 2.2 times increase in wait time. The wait time increase in this case is because the compilation thread waits for the disk thread to service the key management and pointer operations. The wait time is more pronounced in this benchmark compared to Postmark, because kernel compilation has a small I/O component by virtue of its CPU-intensive nature.

5. CASE STUDY: DISK-LEVEL DATA CONSISTENCY

A key challenge in persistent data storage on disk is ensuring the *consistency* of data in the face of crashes. In many cases, on-disk data is unusable unless it conforms to certain software-specific invariants that define its consistency. For example, an on-disk B-Tree with dangling pointers in some of its nodes cannot be used to locate data items. Similarly, in a file system, a directory pointing to invalid or unallocated inodes constitutes a consistency violation.

Given the importance of consistency, most file systems and other software that manage on-disk storage incorporate mechanisms to ensure on-disk consistency. While some techniques involve optimistically updating on-disk state and then *fixing* consistency violations based on a disk scan (e.g., `fsck`), more modern techniques such as journalling [Gifford et al. 1988] or Soft updates [Ganger et al. 2000] involve constraining updates in such a way that consistency is enforced. These mechanisms are quite complex; for example, modern file systems owe a significant portion of their complexity to satisfying this requirement.

This traditional approach to managing consistency entirely at the file system or software is fraught with two key weaknesses. First, the disk system is completely oblivious to the consistency of the data it stores, which constrains the range of functionality it can provide. For example, today's block-based disk systems cannot perform consistent snapshotting of data. Snapshotting is a popular and useful feature in the storage industry, but consistent snapshotting has so far been restricted only to storage systems exporting a richer NFS-like interface [Hitz et al. 1994]. Similarly, modern storage systems perform backup and asynchronous remote mirroring [Ji et al. 2003]; consistency-awareness at the storage level can increase the utility of these techniques.

A second problem with the current approach to consistency management is that every file system and every software layer that manages on-disk data is forced to duplicate the mechanisms needed to enforce consistency. This raises the bar for implementing any disk-resident data structures. Although applications can use generic transactional libraries, it often requires restructuring the application to be aware of transactions and tracking transaction context across concurrent, asynchronous operations. For example, although the journalling block device (JBD) layer in Ext3 provides a transactional interface, the Ext3 codebase had to go through a substantial amount of restructuring to actually use JBD [Tweedie 1998].

To address these problems, we present *ACE-Disk*, an **A**utomatic **C**onsistency **E**nforcing Disk, a disk system that preserves the semantic consistency of stored data. In our approach, the disk system takes responsibility for consistency management, and thus is empowered to provide consistency-aware functionality such as snapshotting. Applications simply inform the disk about the relationship between various blocks that the application already knows about. Specifically, we advocate using a *Type-Safe Disk* (TSD) [Sivathanu et al. 2006], a disk system that is aware of the pointer relationship between blocks, to get consistency, with minimal modifications at the software-level.

Our disk-level consistency mechanism enforces the following constraint: the on-disk version of data should always be consistent. To accomplish this, we need to discover semantically consistent groups of blocks and commit them atomically to the disk when they are written by higher level software such as the file system. All inconsistent block updates should be buffered inside the disk until they become consistent. For example, when a new file is created, the corresponding directory block and the inode block have

to be updated. When just one of the writes arrives at the disk it indicates an inconsistent update. In that case, we need to buffer the update until the second block write also arrives. When both the directory block and inode block writes have arrived at the disk, we need to ensure (at the disk level) that both these blocks are committed atomically to stable storage.

In this section, we describe the main aspects of our disk-level consistency mechanism. First, we discuss some related work. Second, we describe how update dependencies between blocks can be inferred from pointers. Third, we present our enhanced pointer interface that make dependency inference robust. Fourth, we describe the consistency enforcement process a key issue in disk-level consistency enforcement. We finally detail our prototype implementation of the system, and discuss some limitations of pointer-driven consistency.

5.1 Inferring Dependencies from Pointers

Determining semantic relationships between blocks at the disk level requires additional information exchange between the software layer and the disk. Today's block-based disks treat all stored information as opaque data and they do not have knowledge of data semantics. For example, today's disks cannot differentiate between a data and meta-data block in a file system. We leverage the idea of Type-Safe Disks (TSDs) [Sivathanu et al. 2006], to obtain pointer-relationships between blocks as maintained by the higher level software.

Pointers at the disk level not only convey structural information about data items stored on disk, but also they enable the disk to infer dynamic relationships between blocks that get updated. For example, when a new block a is allocated and a pointer is created to it from another block b, both a and b depend on each other. If the system crashes when just one of the blocks is updated, the disk is left in an inconsistent state.

The existing TSD interface consists of primitives for allocation and pointer operations as discussed in Section 3.2. We discuss how each TSD primitive can be used to infer update dependencies.

The allocation primitive internally creates a pointer to the newly allocated block, in the reference block passed. This operation relates two blocks: the newly allocated block and the reference block. Updating one of the blocks alone clearly leaves the system in an inconsistent state; hence these two blocks constitute a dependency constraint and they have to be committed atomically to stable storage.

The pointer creation primitive creates a pointer from any two arbitrary allocated blocks. In this case, the source block *must* be written subsequent to the pointer creation operation to write the new pointer value in it. However, the destination block need not necessarily be written, as the it is a previously allocated block. For example, while creating a new file in the Ext2 file system, a pointer gets created from the directory block to an already allocated inode block that contains the inode of the new file. In this case, both these blocks constitute a dependency. This is because the directory block has to be updated with the new pointer to the inode block, and the inode block has to be updated with valid information about the newly created file. Failure to commit the latter will result in a directory entry pointing to an invalid inode. As a counter example, if we consider a common index-based storage structure, a set of index blocks point to data block. In this case, duplicating an index block for reliability reasons would result in creation of new pointers from the duplicated index block to the existing data blocks. Here only the index block needs to be written and not the data blocks. Therefore, the pointer creation primitive provided by TSD does not convey enough information to decide whether or not the source and destination blocks constitute

a dependency.

A pointer deletion operation deletes an existing pointer from block a to block b. This operation has a special case: if the deleted pointer is the last incoming pointer to block b, we garbage collect b and it can be re-allocated during future allocation requests. In both cases, it is clear that block a has to be written subsequent to this operation for it to reflect the pointer deletion. The destination block b in the case of garbage collection need not be written. However, it does constitute a dependency: b must not be re-allocated until a is written. For example, when the last pointer from an inode block to a data block is deleted during a `truncate` operation, re-allocating the data block to another inode before the old inode is written could result in a state where the old inode points to the contents of a different file. In the normal case of a pointer deletion where garbage collection does not occur, we cannot infer whether the source and destination constitute a dependency for the same reason as explained in the case of pointer creation.

5.2 An Enhanced Pointer Interface

As described in the previous section, the pointer API exported by a TSD do not always convey enough information to make correct inferences in a generic manner. In this work, we fine-tune the TSD API to make it more complete in terms of conveying pointer information.

We introduce the notion of a *sub-block* in a TSD. We use sub-blocks to formalize *allocatable* units inside a block, as maintained by the higher-level software. For example, in Ext2 each inode block can contain several inodes, each of them allocated and freed at the software level. Although formalizing these units in a precise manner requires knowledge about the unit size and offsets inside a block, we just need a rudimentary knowledge of sub-blocks to infer dependencies. For example, to decide whether or not a create or delete pointer operation constitutes a dependency we just need to know if that pointer points to a sub-block. This intuition is based on the fact that, to preserve pointer consistency we need to guarantee two properties: first, no pointer points to unwritten (junk) units, and second, no allocated units become unreachable. In our inference mechanism we make use of additional disk primitives for creating and deleting pointers to sub-blocks. Note that the disk need not track information about sub-blocks, but it just needs to dynamically know sub-block pointer operations by way of explicit primitives. Higher-level software call the respective sub-block primitives while creating and deleting pointers to newly allocated or freed sub-blocks. For example, Ext2 has to call a sub-block pointer creation primitive to create a pointer between a directory block and inode block while creating a file. From this we can infer that the directory and inode blocks form a dependency constraint.

We present an extended pointer interface to TSDs that captures most cases of dependency inferences. In the primitives described below, the parameter t refers to a logical timestamp value for the operation. This is to let the disk know about the temporal ordering of operations as they are issued by the higher level software. The purpose and usage of this parameter is discussed in detail later in this section.

(1) READ($Blockno$): Block read primitive.

(2) WRITE($Blockno$, t): Block write primitive.

(3) ALLOC_BLOCK(Ref, t): Allocates a new block a from the disk-maintained free-block list and creates a pointer to it in Ref. Both Ref and a constitute a write dependency constraint.

(4) CREATE_PTR(Src, $Dest$, t): Creates a new pointer from Src to $Dest$. This primitive does not create any dependency.

(5) DELETE_PTR(Src, $Dest$, t): Deletes an existing pointer from Src to $Dest$. If this is the last incoming pointer to $Dest$, $Dest$ is garbage collected (marked free) and it creates a new dependency between the write of Src and the re-allocation of $Dest$.

(6) MOVE_PTR(Src, $Dest$, $Newsrc$, t): Moves the source block of an existing pointer from Src to $Newsrc$. This operation results in creation of a new dependency for the writes of Src and $Newsrc$. This primitive is useful for handle cases such as a `rename` operation in a file system, or a B-tree node split where pointers need to be moved from one block to another.

(7) ALLOC_SUB_BLOCK(Ref, $Target$, t): Creates a new pointer between block Ref and block $Target$. $Target$ is a block that contains multiple allocatable software-level structures. This primitive is called when a software-level structure in $Target$ is allocated. This disk does not track these structures. This creates a new write dependency between Ref and $Target$. The disk differentiates this primitive from the CREATE_PTR primitive only to infer dependencies.

(8) FREE_SUB_BLOCK_PTR(Ref, $Target$, t): Deletes an existing pointer between Ref and $Target$. $Target$ is a block that contains multiple allocatable software-level structures. This primitive is called when a software-level structure in $Target$ is freed. If this operation deletes the last incoming pointer to block $Target$, $Target$ is garbage collected and a new dependency is created between Ref update and re-allocation of $Target$. If the pointer deleted is not the last incoming pointer to $Target$, a new dependency is created for the update of Ref and $Target$.

5.3 Consistency Enforcement

In this Section we detail how an ACE-disk guarantees consistent data commits to stable storage. Figure 10 shows the overall architecture of an ACE-disk.

An ACE-disk consists of five main components: (1) *dependency buffer*, a buffer layer made of high-speed memory where inconsistent block updates are buffered until the corresponding dependency becomes consistent; (2)*buffer swap space*, a swap area in the disk which is used to swap out inconsistent buffer data when the cache is full; (3) *journal space*, an area on disk which is used to ensure atomic update of resolved dependencies; (4) *group manager*, which tracks the pointer operations and constructs dependencies; *group index* a data-structure used by the group manager to store disjoint dependencies and the blocks affected by each of those dependencies. The buffer layer acts both as a read and write cache, and gets invalidated during power down of the disk. All inconsistent block updates are buffered in the cache to ensure that the state of data stored in place is always consistent. The swap space is used when the number of inconsistent blocks exceed the size of the high speed buffer memory.

When an ACE-disk infers a dependency during a pointer operation, it associates a *group* object with that dependency. This group object contains information about the set of blocks that are affected by that dependency. We use the terms *group object* and *dependency group* interchangeably in the rest of the report to refer to a list of blocks that needs to be committed atomically to stable storage to ensure consistency. A *group entry* refers to a member of a group which contains a block number and the time at which it was added. When a block is written *after* it is added to a dependency group, the corresponding group entry for that

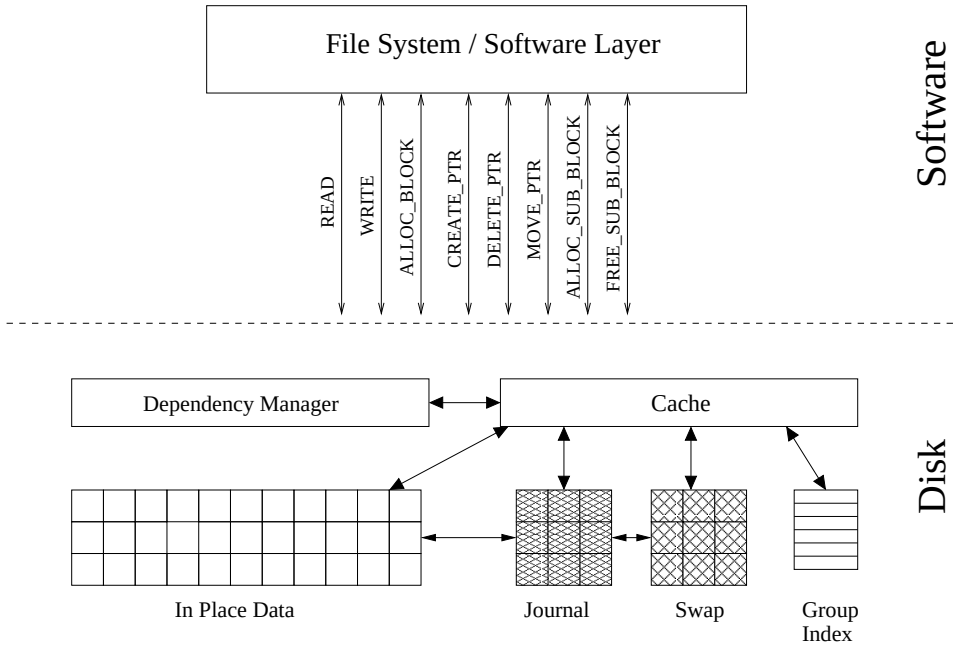


Fig. 10. *Architecture of an ACE-disk*

block is marked "ready." When all entries in a dependency group are ready, the group is said to be *resolved*, and all blocks associated with it can be committed atomically to the disk.

In a simple case, when the first pointer operation happens in a disk causing a dependency creation between two blocks a and b, a new dependency group G is created and both the blocks are added to it. When write requests for both a and b have arrived at the disk, the dependency group G is said to be *resolved* and all the blocks in G can be committed atomically to the disk. However, if another pointer operation happens before G is resolved introducing a dependency between blocks b and c, the operation *extends* the existing dependency group. This is because, one of the blocks in the new dependency (block b) is already part of an existing dependency. Thus, in this scenario block c should be added to group G as well. Therefore, whenever there is a new dependency introduced between any two blocks x and y by way of a pointer operation, one of the following three actions are taken:

(1) If both x and y are not part of any existing dependencies, a new dependency group is created and x and y are added to it.

(2) If only one of x or y is associated with an existing dependency group G, then both blocks are associated with G and are marked "not ready."

(3) If both x and y are already associated with the same group G, then no group action needs to be taken. However, the entries in the group pertaining to blocks x and y have to be marked "not ready" as a new constraint is added between the two blocks.

(4) If both x and y are associated with different groups G_1 and G_2, then G_1 and G_2 are

merged, and the entries for x and y are marked "not ready"

As pointer operations construct dependencies between blocks, higher-level software must ensure that the pointer management primitives are issued to the disk *before* the source and destination blocks are updated. This constraint is implicitly enforced for the block allocation primitive as a block cannot be updated before it is allocated. However for the pointer creation and deletion primitives, higher-level software has to ensure that it follows this ordering rule. For example, when a `create` happens in Ext2, the sub-block pointer creation primitive has to be issued for the directory and the inode blocks before the contents of the blocks are updated.

Temporal Ordering of Operations. ACE-disk's consistency mechanism relies on the temporal relationships between operations seen at the disk level. For example, an entry in a dependency group is marked ready when a write arrives after the dependency creation. However, in today's modern operating systems and disks, operations can be re-ordered at any level. For example, file systems today predominantly perform asynchronous I/O where block writes are buffered at the software level and are flushed to the disk in regular intervals of time. Moreover, modern disk device drivers re-order or merge disk requests before issuing to the disk for performance reasons. These factors make the temporal ordering of operations that the disk sees completely different from the order that the higher-level software issued. Therefore, unless additional ordering information is communicated from the software-level, the disk cannot obtain the precise temporal order of operations.

ACE-disk solves this problem by introducing two constraints on the operations: (a) all pointer primitives take place synchronously and (b) all operations have associated logical timestamps. These two constraints enable the disk to obtain precise temporal ordering of the operations. Although synchronous pointer operations may affect performance, it is mitigated by the fact that these operations do not result in block I/O inside the disk, in the critical path. Timestamps in this case are logical. For example they can be a monotonically increasing sequence number. Whenever higher-level software issues a pointer operation, it has to pass a sequence number along with it. Similarly when the in-memory copy of a disk block is updated by the software, a sequence number has to be associated with the buffer for that block. Whenever a pointer operation introduces a dependency, its sequence number is associated with the corresponding group entries. The entries are marked ready only when a subsequent write arrives with sequence number greater than the stored one. Note that introducing sequence numbers with block I/O operations is simple—we have modified the Linux kernel to support sequence numbers along with buffers whenever they are dirtied. This modification was trivial and required changing just 50 lines of code.

When a dependency group is resolved all blocks in the group has to be committed in place atomically. A power failure while committing a dependency group should not leave the in place data in an inconsistent state. ACE-disk uses a logging mechanism to ensure this. All blocks in a resolved groups are first written to a log and synced with a commit identifier before the in place commit happens. The log is discarded when the in place commit is complete. After a crash, an ACE-disk checks the log for valid group data and replays them. The log contains separate journals for each dependency group and hence each of them are replayed after the crash to bring the system to a consistent state.

5.4 Bounding Commit Interval

The amount of data lost during a crash depends on the interval between the instant a block write arrives at the disk and the time when it is actually committed to stable storage. In an ACE-disk, inconsistent block data gets buffered until the entire dependency group is resolved. ACE-disk's mechanism of managing dependency groups allow extending a group whenever pointer operations happen from or to a member of the group. Thus, during normal operation, a dependency group could potentially get extended repeatedly during a continuous workload that performs pointer operations. For example, in Ext2, for a recursive directory creation workload, the entire working set would form part of the same dependency group as all blocks branch out from the inode of the root directory. Moreover, as pointer operations always precede the block write operations, a dependency group could never get resolved for a continuous workload. This is because before the time when all blocks in a group are marked ready, the group could be extended several times with new blocks or new dependencies for the existing blocks. This results in two problems. First, large amounts of data may get lost in the event of a crash, although the on-disk state is consistent. Second, excessively long dependency groups require buffering of a large number of blocks and hence impose onerous space requirements.

Bounding the interval between dependency commits is challenging particularly at the disk level because the disk has no knowledge about intermediate versions of block data that are known to the higher-level software. This is because most higher-level software buffer writes and hence the versions of block data that reach the disk could be a small subset of total number of versions that the software knows about. For example, if a file is created in Ext2, an inode block is modified. Before the inode block write is issued to the disk, if another file is created whose inode is in the same block, the disk sees only the version of the block updated with both inodes. Therefore, the disk cannot spawn a new dependency group during a pointer operation for a block, when the existing group containing a block has reached a time threshold.

Blocking pointer operations at the disk level until an existing dependency is committed could be a solution to the bounding problem, but requires radical modifications to the higher-level software to support it. This is because software such as file systems perform locking of data-structures at an operation level. When a pointer operation blocks, the file system could sleep after grabbing a lock on the data-structure which reside on a block that needs to be committed for some dependency to resolve. This could result in a deadlock as the block containing the data-structure cannot be committed until the operation in execution completes.

An ACE disk solves this problem by having new error modes for pointer creation operations. The allocation and pointer management primitives could optionally return one of the following errors to the higher-level software: SYNC_BOTH, SYNC_SRC, or SYNC_DEST. As the names indicate, the disk can fail a pointer operation and choose to request the higher level software to write the source, destination, or both blocks associated with that operation. Upon receiving one of these errors the software should issue a write of the current version of the corresponding blocks, and then retry the pointer operation. At the disk level, whenever a dependency group is unresolved beyond a time threshold it is *frozen*. Whenever new dependencies are created for a block that is already part of a frozen group and in an "not ready" state, the disk returns one of three errors mentioned above, depending on whether the block is the source, destination, or when both the source and destination blocks

exist in frozen groups in "not ready" state. This way of forcing the software to commit the intermediate version of the data helps the disk to spawn new dependency groups for blocks that are already ready in a frozen group. An ACE-disk ensures that at a block is never part of more than two groups at a time, the older of which is frozen. This is done by ensuring that a group is not frozen until all blocks in the group are not part of any other frozen group. This method ensures commit of dependency groups in tune with the block write interval of the higher level software. We verified the correctness of our bounding solution by implementing this in the Ext2 file system. Each every case, the commit interval of the dependency groups were in tune with that of the software level write-back interval.

5.5 Implementation

We implemented a prototype ACE-disk as a pseudo-device driver in the Linux kernel 2.6.15 that stacks on top of an existing disk block driver. The pseudo device driver layer receives all block requests, and redirects the common read and write requests to the lower level device driver after the required processing. The additional primitives required for operations such as block allocation and pointer management are implemented as driver `ioctl`s.

To enable sequence numbers with block I/O requests, we added a new field to the buffer header object and the `request` token object in the Linux kernel. Whenever a buffer is marked dirty, we generate a sequence number and update it in the buffer header. When a write is issued for a buffer, the sequence number is carried over to the `request` object and hence available to the ACE-disk pseudo-device driver. Sequence numbers are generated by an atomic increment of a counter value. The same counter value is used during pointer operations and modifying buffers. Our prototype ACE-disk contained 6,900 lines of kernel code of which 3,060 lines of code were reused from the existing TSD prototype.

5.6 Limitations of Pointer-driven Consistency

While the update dependency information conveyed by pointers is quite rich and as we show, sufficient to enforce consistency, it has some limitations when compared to the more general notion of transactional consistency. Specifically, the dependency information conveyed by pointers is limited to a pair of blocks; e.g., if a pointer is created between two blocks, the two blocks will be updated atomically. However, our mechanism cannot support atomic commits of an arbitrary group of blocks. For example, on creation of a new directory (i.e., mkdir) in ext2, a pointer is created from the parent directory block to the inode of the child directory, and the inode initialized. Then a new block is allocated for the child directory and a pointer created between the child inode and the child directory's new data block. With a transactional system, these three blocks will be committed atomically. But in our case, the first pointer creation and the initialized inode could be committed before the second pointer creation. As a result, a directory inode may end up with a state where it has no blocks at all, which is an apparent violation of consistency.

However, we argue that this consistency problem falls under a class of *online-patchable* consistency violations. For example, just by looking at the initialized directory inode with no pointers, it is unambiguous that a crash happened just before the new directory's block got allocated, so it is safe to immediately allocate a new block for the directory and assign it to the inode. Note that in contrast, a more "real" consistency problem would be a directory pointing to the wrong inode, perhaps a regular file inode, where it is not obvious what the correct state should be. Pointer consistency could lead to such transient online-patchable consistency violations the violation is readily and unambiguously identifiable and the fix

for that is obvious as well. Most importantly, the fix to such a violation is *local*, in that it does not require looking at the global state of the file system. We believe that the pointer-derived consistency semantics is thus a useful and simpler counterpart to the more general transactional consistency.

5.7 Evaluation

We evaluated the performance of our prototype ACE-disk using Ext2ACE. We ran both a general purpose workload and a micro-benchmarks on our implementation and compared it with a regular Ext2 and Ext3 file systems running on a normal disk. We compared our system with Ext3 because it is a journalling file system that provides similar consistency guarantees as ACE-disk at the software level. For all benchmarks we used Ext3 in its default journalling mode (ordered writes mode). In this mode file meta-data alone is journalled and it is written to the journal only after the corresponding data blocks are written directly in place.

For all benchmarks we included the file system unmount time in our calculation. This is because ACE-disk commits dependency groups asynchronously using separate kernel threads, and a file system unmount procedure blocks until all outstanding threads have completed their commit operation. This is relevant even for normal Ext2 and Ext3 as they commit all outstanding dirty data during an unmount.

Postmark Results. We configured Postmark to create 30,000 files whose sizes ranging from 512 bytes to 10 KB, and perform 250,000 operations in 200 directories. This workload particularly stresses the ACE-disk as a large number of dependencies get created and resolved during the meta-data operations. The time taken for the Postmark benchmark for Ext2, Ext3, and Ext2ACE are shown in Figure 11(a).

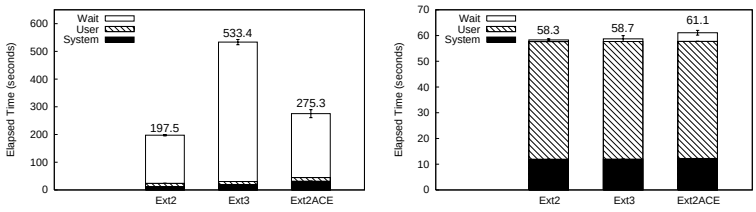


Fig. 11. *(a) Postmark and (b) OpenSSH compile Results for ACE-Disk*

Ext2ACE on top of ACE-disk had an elapsed time overhead of 40% compared to regular Ext2 on a normal disk. Although the system time increase is 2.6 times relatively, this has not contributed much to the elapsed time overhead. As mentioned earlier, this overhead is because of dependency tracking during every block write and pointer operations. The wait time increase (32%) is predominantly because all blocks are written out twice in the case of an ACE-disk to ensure atomic commits of dependency groups. All block data is written out to the journal first and after the journal is synced, in-place commits happen. Ext3 ran almost twice as slow as Ext2 because of its ordered journalling mode. Ext2ACE is faster than Ext3 in this case because ACE-disk journals both data and meta-data blocks and for a small file workload such as Postmark, random writes get converted to sequential ones. The in-place commit of data in ACE-disk happens in an asynchronous manner.

Compile Benchmark Results. To simulate a relatively CPU-intensive user workload, we compiled the OpenSSH source code. We used OpenSSH version 4.5, and analyzed the overheads of Ext3 and Ext2ACE for the `untar`, `configure`, and `make` stages combined. These operations in combination constitute a significant amount of CPU and I/O operations. The results for OpenSSH compilation is shown in Figure 11(b).

The times taken by Ext2 and Ext3 for the compilation workload are almost similar. This is because this is a mostly CPU-intensive workload. Ext2ACE had an elapsed time overhead of 5% compared to Ext2 and Ext3. This is because of the increase in wait time (1 sec vs. 3.4 secs). The increase in wait time is caused by the CPU context switches between the main compilation process and the asynchronous dependency commit threads of ACE-disk. Since this is a CPU-intensive workload, the context switch time is more pronounced than Postmark. In a real environment, as the dependency commits are performed inside the disk, this context switch overhead would not be seen. The system time overhead is not significant for Ext2ACE in this case because there are relatively few I/O operations that require processing to track dependencies.

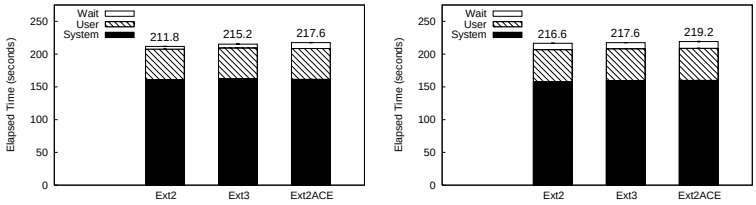


Fig. 12. *(a) Create and (b) Unlink micro-Benchmark Results for ACE-Disk*

Micro-Benchmarks. We ran two micro-benchmarks to obtain the overheads of the `create` and `unlink` file system operations. We evaluated these two operations because both of them exercise the ACE-disk's dependency trackers and consistency enforcement mechanism. For the create workload, we created 500 directories with 1,000 files each totaling to 500,000 files. For the unlink workload, we removed all created files and directories. The results of the `create` and `unlink` workloads are shown in Figures 12(a) and 12(b), respectively.

For the `create` workload, Ext2ACE had an overhead 2.7% compared to Ext2. This is mostly caused by the increase in wait time due to the additional I/O operations writing out block data twice for ensuring atomicity in block commits. For the `unlink` workload the results of Ext2ACE is similar to Ext2 and Ext3 as `unlink` results in smaller number of writes than creates, because freed blocks are not written to the disk.

Overall ACE-disks have small overheads for normal user workloads. When the workload is highly I/O-intensive, more information needs to be tracked by the disk to manage dependencies. This results in more CPU time which is mitigated by the fact that the disk uses its own isolated CPU in a real environment.

6. CONTEXT-AWARE I/O INFRASTRUCTURE

In this section, we present the concept of *Context-Aware I/O* (CAIO), a simple and generic way for applications to convey arbitrary information about their I/O behavior and relation-

ships, without worrying about how the information will be used by the storage stack. In CAIO, an application-level *context* is propagated along with an I/O operation across the entire storage stack, in an end-to-end fashion. An application-level context is represented by one or more *context identifiers*. For example, a database application can have a unique identifier that it can propagate along with every I/O it generates, such that any storage layer can easily group all I/O generated by the database application. This also enables the lower layers of the storage stack to associate the data corresponding to the I/O with higher-level contexts and easily track the application's working-set.

In addition to working-set identification, application contexts also enable a new class of functionality that uses application-I/O relationships, such as easy and flexible performance isolation in large-scale distributed storage, and access-pattern aware caching and prefetching within the storage hardware.

To make CAIO a generic framework, we decouple the *generation* of application-level information from how the information is *used* within the storage stack. Most hint-based proposals to address the problem of information-gap in the past have tied these together. For example, in hint-based prefetching systems, the application provides hints of its future access, but the hints are specifically designed with prefetching in mind. The problem with such function-specific hints is that they require coordination and agreement between the layers involved. In a multi-vendor setup, such coordination translates into industry-wide consensus on the interface, a standardization process that takes years. In addition, such an approach cannot scale in an end-to-end manner to the multi-layered storage stacks that we have today.

Decoupling the generator and consumer of the context information leads to an interesting challenge: when the application could conceivably use more than one possible granularity of grouping I/O, how can it decide which one to use while being oblivious to how the grouping is interpreted by the lower level? For example, a database application can group the I/O requests it generates based on the database user, session, transaction, or query on behalf of which the I/O is issued; but the lower layers are oblivious to the granularity of the context. To solve this issue, contexts in CAIO are *hierarchical*. With hierarchical contexts, higher layers can encode multiple granularities of grouping, and the lower layers can decide which granularity is the best for the particular functionality that they provide.

Even in a hierarchical context, individual levels in the hierarchy remain completely opaque to the storage stack. For implementing functionality that needs more information about what these levels in the context mean, contexts can be annotated *offline* at any specific layer. In such cases, CAIO contexts will be used only as naming-identifiers to associate higher-level semantics.

We illustrate the generality and power of the context abstraction by prototyping and evaluating two case studies. Our first case study is an automatic working set identifier, *WorkSIDE*, which operates at the block-based storage hardware layer. WorkSIDE automatically tracks the data working set required for an application context to run to completion. WorkSIDE correlates contexts with the I/O and the corresponding data they access, thus obtaining a complete view of the entire set of data items that the particular application context requires. This working set can then be preloaded as appropriate in order to improve performance and availability, or to enable power optimizations. The second case study is a context-aware cache-placement algorithm within the disk that automatically tracks which application-level contexts exhibit sequential streaming access pattern and avoids caching

requests with that context. We demonstrate the usefulness of both of our case-studies using prototype implementations we built for the Linux kernel, and evaluate various workloads.

The rest of this section is organized as follows. In Section 6.1 we discuss the utility of CAIO by presenting a few potential applications. In Section 6.2 we present a taxonomy of the various kinds of contexts in storage. We detail how we generalize the CAIO interface in Section 6.3. In Section 6.4, we describe CAIO design and application support.

6.1 The Utility of Context-Aware I/O

In this section we describe several usage scenarios that motivate tracking context information in the different layers of the storage stack. Many of these utilities cannot be implemented effectively without explicitly propagating application-level contexts. In Sections 7 and 8, we demonstrate our implementation of the first two usage scenarios described below.

Working-set Aware Features. Identifying working sets of data for individual applications at the lower layers of the storage stack, enables interesting functionality such as application-aware prefetching [Patterson et al. 1995], power-savings [Zhu et al. 2005; Weddle et al. 2007], selective recovery of failed hardware [Magoutis et al. 2007], and improved data availability [Sivathanu et al. 2004]. We describe our implementation of a disk-level working-set identifier and its usefulness in detail under Section 7.

Adaptive Caching and Prefetching. The efficacy of caching and prefetching depends on the ability to identify access patterns. Context can enable caching and prefetching mechanisms to adapt their policies based on access patterns. Section 8 describes our implementation of a context-aware disk-level caching mechanism.

Application-Aware Performance Isolation. Scheduling algorithms at different levels of the storage stack can leverage application-level contexts in scheduling decisions. For example, fair share disk schedulers can enforce fairness based on higher level logical tasks as against OS processes. Application-based resource isolation has been previously explored in the context of a single OS in Resource Containers [Banga et al. 1999]. Contexts can enable flexible resource isolation in an end-to-end fashion even in distributed storage.

Optimized Data Layout. File systems can use higher level contexts as hints for optimal data placement on disk. Co-locating files and directories created in the same context could be beneficial under certain scenarios to achieve better spatial locality during reads.

Improved Accounting. Context information associated with I/O operations can greatly help in I/O trace analysis. Trace analysis for resource consumption can be more accurate when it makes use of logical contexts pertaining to precise higher-level tasks. Contexts can also provide valuable hints about the dependencies of I/O operations and the causal relationships between them, for trace-based intrusion detection systems [King and Chen 2003].

6.2 Context Types

Context in storage is quite useful as seen from the kind of functionality it enables (described in Section 6.1). We now define *context* as follows: *A context in storage is a reference or identification used to group, on some basis, several I/O operations or data*.

We now describe the types of contexts that are relevant to storage.

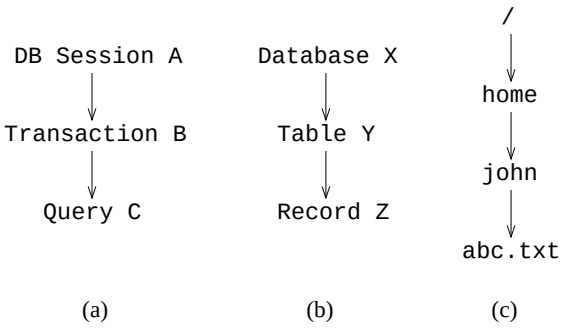


Fig. 13. *Examples of how hierarchical contexts can be constructed. (a) shows an access-bound context hierarchy. (b) and (c) show data-bound context hierarchies.*

Data-Bound vs. Access-Bound. The two primary entities in storage are (a) data, and (b) I/O operations on data. Context in storage is mainly used for grouping several such data items or I/O operations. Therefore we classify context in storage broadly into two types: data-bound and access bound.

A context is said to be *data-bound* if it can be used to group several data items stored on disk, based on some metric. This grouping is independent of the way the data is accessed. For example, a data-bound context can group all blocks belonging to the same database table or file. Data-bound contexts can group data based on arbitrary criteria such as logical abstractions (files, directories, database tables, etc.), owning application or user, security domains, and so on. Data-bound contexts can be used to communicate higher-level data-structures to the disk, and enable functionality such as fault-isolated placement in RAID [Sivathanu et al. 2004].

Note that the notion of data-bound contexts is similar in concept with other abstractions such as type-aware storage (Section 3) or object-based interface [Mesnier et al. 2003]. These other abstractions can be used as an alternative to data-bound contexts.

Access-bound contexts relate operations rather than the data pertaining to them. For example, an access-bound context can group all block write operations resulting from a single database query. Access-bound contexts enable new functionality that solely depend on the characteristics of individual I/O requests. The caching and prefetching functionality described in Section 6.1 requires access-bound contexts.

Figure 13 shows a few examples of context hierarchies. Figure 13(a) shows a possible access-bound hierarchy for a database application. Figures 13(b) and 13(c) show data-bound context hierarchies that communicate data abstractions.

Repeatable vs. Non-Repeatable. The lifetime of a context identifier is defined by the application that generates it. When a single context identifier is used every time to refer to a particular logical context, we call it a *repeatable* context. For example, when a context is used to group files within an access-control domain, the same identifier has to be reused every time when operations are performed on that domain. Applications have to generate such contexts using a deterministic method and may maintain persistent states to track contexts.

Non-repeatable contexts have transient identifiers. For example, if a `pid` is used as a context identifier to group I/O operations generated by a particular program, every time

the program runs, the identifier becomes different, although the logical context remains the same. Non-repeatable contexts do not require any state to be maintained at the application-level.

6.3 Generalizing the Interface

In this section, we describe how we can cope with arbitrary context generation process at the application-level, and achieve independence between the generation and usage of application-contexts. We also describe how lower layers of the storage stack can extend contexts or correlate across different context types.

Hierarchical Contexts. To achieve generality in the CAIO interface, the context generation process at the application-level must not make any assumptions about how the lower layers use the context. However, at the application-level, there may be several different ways to generate a context, each useful for different kinds of functionality at the lower layers. A single application-wide context identifier can be used to easily group all data required by the application, whereas more fine-grained context identifiers within an application help communicate different streams of I/O requests generated by sub-components within same application. For example, a single DBMS-wide context can be used to group all I/O and data that the DBMS manages. This enables functionality such as working-set identification for the entire DBMS. On the other hand, a per database session-level context can be used for easy performance isolation between database user sessions. We use the term *context granularity* to refer to the different possible ways to generate contexts within an application.

Therefore, for generalizing the interface without hampering the kind of functionality it enables, we evolve a context scheme where the application can encode all possible granularities as a single context, passing down *context hierarchies* (for access-bound and data-bound) rather than a single identifier. For example, a DBMS can generate access-bound contexts in granularities such as sessions, transactions, and individual queries, and data-bound contexts in granularities such as databases, tables, and records.

Lower layers of the storage stack can use hierarchical contexts without making assumptions about what each of the levels in the hierarchy mean. For example, a caching layer that wants to classify some context to exclude caching (e.g., sequential contexts) can track the statistics on sequentiality at each level of the context hierarchy, and then choose the highest level that exhibits homogeneity in the access pattern. Depending on the specific behavior the layer is looking at (e.g., sequentiality, correlated access of the same pieces of data), the definition of homogeneity changes. Hierarchical contexts enable decoupling the application from worrying about which behavioral properties the lower layers are interested in; instead the application just conveys its state, and the lower layers make their independent decisions on the notion of homogeneity they care about, based on the layers' own per-context statistics.

Note that for a context hierarchy chain in CAIO to be meaningful, every context in the chain should qualify a logical subset of the access or data domain qualified by its parent context. For example, a per-query context identifier can be a child of the transaction identifier in which the query is a part. However, a context identifier that qualifies the class of *all* `select` queries in a DBMS cannot be a child of any particular transaction identifier, as `select` queries can be part of any transaction.

Annotating Contexts with Semantics. Certain functionality may require more information about what each level in the hierarchy means, at some specific layer in the storage stack. For example, a context-based proportional-share disk scheduler needs share proportions to be associated with levels in the context hierarchy. For this purpose *offline* mechanisms can be used to annotate context identifiers with functionality specific information. For example, applications can co-ordinate with a specific file system through offline *ioctl*s to associate locality hints with stored context identifiers. Note that these annotations are not part of the CAIO infrastructure, but can be done separately between any two layers that needs to coordinate to implement a specific functionality. In the example of a proportional-share disk scheduler, the application and the disk scheduler need to co-ordinate offline to annotate context levels with share proportions.

Context Transformation. With hierarchical contexts, any layer in the storage stack can add new levels to the context chain, as long as the subset invariant is preserved. For access-bound contexts, the subset relationship is maintained as an operation propagates from top to bottom. For example, a *select* query generated from a database gets transformed into one or more file read operations at the file system, and then further transformed into several block read operations at the device driver or the disk level. Therefore, any layer in the stack can add new levels to communicate grouping of sub-operations at their level.

However, for data-bound contexts, subset relationship is harder to ensure across layers. This is because the data abstractions used by higher layers in many cases are not supersets of the lower level abstractions. For example, an application can store several B-trees within a single file, and hence there is no subset relationship between the abstractions used by this application and that of the file system. Therefore generic transformation of data-bound contexts across layers is harder to achieve; but lower layers can associate new data-bound context hierarchies with I/O, if the application does not pass a data-bound context. We impose a constraint that intermediate layers should not add new levels to data-bound contexts, unless the higher-level layer did not specify a context of its own.

Correlating Across Context Types. Data-bound and access-bound contexts passed by the application can be completely independent of each other and need not necessarily indicate association between the operation and the data it operates. This makes generation of contexts at the application-level much less complicated. However, lower layers that use these contexts can maintain their own history information of contexts, and correlating data-bound and access-bound contexts. Correlating context types enables useful functionality. For example, identifying the working set of data accessed by an access-bound context can be useful for implementing interesting optimizations as described in our first case-study detailed under Section 7.

6.4 CAIO Design

End-to-end association of context with I/O requires passing application-generated context with every I/O operation throughout its lifetime. We evolve a framework through which context can be passed from an application all the way down to the storage hardware (e.g., a disk). In this section, we describe the changes required to the storage stack and user applications, to support contexts.

We propagate context in the storage stack by means of *context objects*. A context object contains upto two context chains, one each for data-bound and access-bound types. These context types are based on the discussion under Section 6.2. Context objects also carry

information about the repeatability of the context chains. Repeatability is at the granularity of an entire chain and not the individual context identifiers within a chain. The structure of a context object is shown in Listing 1.

```
struct caio_context {
    int data_bound[MAX_DATA_LEVELS];
    int access_bound[MAX_ACCESS_LEVELS];
    short data_levels;
    short access_levels;
    int flags;
};
```

Listing 1. Structure of a context object. The fields data_levels and access_levels indicate the number of levels in the data and access-bound context chains. Flags contain information about repeatability and inheritance properties (Section 6.4) for the context.

Associating Contexts with I/O. The CAIO framework contains a user library that exports routines to construct context objects and add new levels of hierarchy to existing context objects. User applications can generate context objects through these routines and associate them with I/O operations. Our framework provides three different ways for user applications to associate contexts with I/O operations. They are, (a) an extended system call interface (b) group contexts and (c) context inheritance. We detail each of these mechanisms below.

An Extended System Call Interface. We have an extended system call interface that passes context objects along with storage primitives such as `open`, `read`, `write`, `unlink`, etc. Each of these I/O system calls include an additional argument for the context object. The framework also includes a wrapper library for user applications to call these new system calls. Listing 2 shows an usage scenario for the extended system call interface. Note that when there is a single context object that needs to be passed for several system calls, *group contexts* can be used for better performance, as described below.

Group Contexts. For applications that need to perform a several I/O operations with a single context object, we provide a new system call for setting and unsetting contexts into the kernel. The scope of this association is just the specific thread of execution. Therefore applications can first set a context and then issue any number of regular I/O system calls (such as `open` or `read`), and the corresponding context object will be associated with every operation.

Context Inheritance. To support easy usage of contexts in cases where the smallest granularity is a process, our framework includes a context inheritance mechanism using which any process can set an *inheritable context* into the kernel. All child processes and threads of such a process will then inherit the same context hierarchy. We developed this feature so that there would be no modifications required to applications whose lowest context granularity is a process. For example, if a project compilation task requires several applications such as `gcc`, `ld`, `binutils` etc., the entire compilation task can be run through a shell that has an inheritable context set, instead of modifying every application to pass contexts.

```
int fd; char buf[128];
struct caio_context *context;

/* Allocates and sets top-level databound
 * and accessbound identifers as 1 */
context = caio_create_context(1, 1);

/* Adds a new level to the access/data
 * hierarchy with identifier 2 */
caio_add_level(context, 2, 2);

/* CAIO system call interface */
fd = caio_open("/home/joe/abc.txt",
          O_RDONLY, &context);
err = caio_read(fd, buf, 128, &context);
caio_close(fd, &context);
```

Listing 2. Passing contexts from the user-level using the CAIO extended system call interface. Note that in this case group context (described in Section 6.4) can be used as well, because a single context object is used for all calls.

Context Propagation. In CAIO, each layer receives contexts from the layer above and passes it to the layer below after using them if applicable. Note that a single operation at a particular layer could translate into multiple operations in the layers below. For example, a file create operation at the file system level could result in multiple block write requests to the device driver. Therefore it is each layer's responsibility to propagate context objects appropriately to the layer below. In cases where there are more virtualization layers such as software RAID or logical volume managers (LVMs), such layers should be aware of contexts and propagate them below. Any layer can choose to store contexts in its own structures for its needs, before passing them down.

Hardware Interface Extensions. To propagate contexts end-to-end, we extend storage hardware interfaces to pass generic context objects along with every I/O request. For example, the SCSI/IDE `read` and `write` primitives take context objects. There are a number of proposals in the past that suggest interface extensions to disk systems for communicating higher-level semantic information [de Jonge et al. 2003; Mesnier et al. 2003; Sivathanu et al. 2006; MacCormick et al. 2004]. We believe that the generality of the CAIO interface would make it easier for disk vendors to adopt.

Dealing with Operation coalescence. Multiple logically independent I/O operations may be coalesced into one at any layer in the stack. For example, multiple file write operations to the contents of the same file block could result in a single block I/O at the disk level due to write buffering. To handle such cases, we support multiple context objects to be associated with a single lower level I/O. Layers that receive these contexts must process them one by one as if they were from different I/O operations.

Storing Contexts. Repeatable contexts may need to be stored by layers to implement optimizations that involve tracking context history, or correlating different context types. We developed a `context-store` in-memory data-structure as part of our framework to enable easy storage of context hierarchies at any layer of the storage stack. A context store manages context hierarchy in a tree structure in which each node represents a context identifier of a specific level in the hierarchy identified by its depth in the tree. Each tree node also includes a *private data* field where information about that specific chain can be stored. The context-store structure provides primitives for common operations such as adding new chains and updating private data.

6.5 Linux Implementation

We implemented our CAIO framework in the Linux kernel 2.6.15. We added new system calls for context-aware file I/O operations and implemented a user-level library for applications to easily use the new system call interface. The new system calls allowed context objects to be passed with `open`, `read`, `write`, `pread`, `pwrite`, `close`, `mkdir`, `unlink`, `rmdir` and `readdir` operations. We modified the following objects to add a new field to store contexts. (a) `task_struct` which represents a running process or thread. (b) `buffer_head` which represents a block buffer in memory. (c) `bio` which represents an I/O to a block device. The `buffer_head` and `bio` objects can optionally contain a list of contexts during operation coalescence.

We implemented the new system calls as wrappers to the unmodified system call handlers for the operations. The wrapper system calls set the context object in the `current` task object before calling the unmodified handlers. Note that the wrapper calls unset the context upon completion of a system call, so that the scope of a passed context would be just that system call. The different layers in the OS that service the I/O operation use the context object from the `current` task object and propagate it to the corresponding `buffer_head` and `bio` objects appropriately. As the `task_struct` object is unique to a particular process or thread, this method works for multi-process workloads as well.

For group contexts, we added a new system call which assigns or removes the corresponding context in the current `task_struct` object. For inheritable contexts, we modified the `fork` system call to copy the context object of the parent, to the forked process. We also implemented the context-store data-structure as part of the kernel so that any layer such as the file system or device driver can maintain its own store.

Overall, the modifications required to implement the CAIO framework were small. We added only 350 lines of new kernel code and 150 lines of user-level code.

Application Support. The method of generating contexts at the application-level depends on specific application architectures. In general, if an application can classify its activities into distinct logical tasks, and (or) if it can group data it uses based on some criteria, it can generate contexts in a meaningful manner. Based on the kind of application, the granularity and type of contexts it can generate can vary. Some low level applications such as Unix utilities (e.g., `ls`, `cat`, etc.) can just provide an interface to the caller to pass contexts (e,g., command line arguments). We have modified some basic utility programs such as `cp`, `cat`, and `ls` to accept contexts as command line arguments. This enables a higher level caller application (e.g., a shell script) to group all its operations under the same context.

Context-Aware MySql. We have modified the MySql DBMS [MySQL AB 2005] with InnoDB [InnoDB 2007] as the storage engine, to generate and propagate contexts at various granularities. MySql has the notion of database client connections which can obtain service from the DBMS. Each client connection gets serviced by a separate MySql thread, and can run several transactions and queries. We modified MySql to pass contexts at three granularities in the form of a hierarchy: connection-level, transaction-level, and a single query-level. Overall the modifications required to propagate contexts across the various layers of MySql and InnoDB were simple. We added only 30 lines of new code and modified 345 lines of existing code, mostly for passing an additional argument for a number of functions. We use our Context-Aware MySql as an application to evaluate our framework and some of the case-studies described in Sections 7 and 8.

6.6 Evaluation

We evaluated the overheads associated with passing context objects across the storage stack for all file system operations. In this section we first describe our test setup and the details of the experiments we ran. Note that the setup described in this section applies to all our benchmarks presented under Sections 7 and 8 as well.

We conducted all tests on a 2.8GHz Xeon with 1GB RAM, and a 74GB, 10Krpm, Ultra-320 SCSI disk. We used Fedora Core 6, running a Linux 2.6.15 kernel. To ensure a cold cache, we unmounted all involved file systems between each test. We ran all tests at least five times and computed 95% confidence intervals for the mean elapsed, system, user, and wait times using the Student-t distribution. In each case, the half-widths of the intervals were less than 5% of the mean.

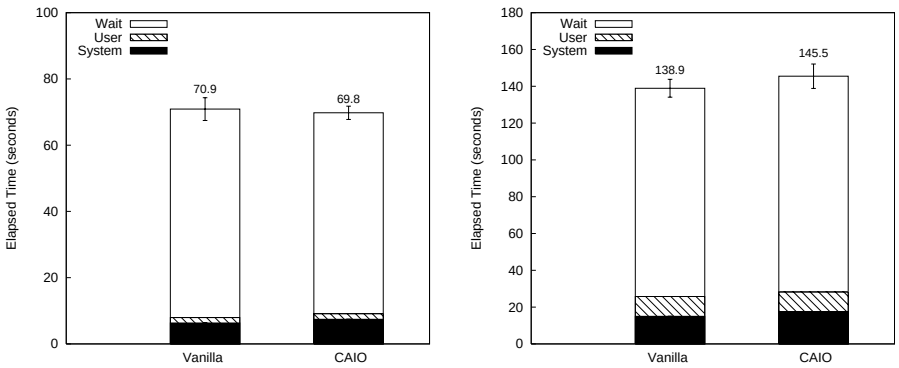
Experiments. In this section we describe the set of experiments and their configurations that we used for evaluating the CAIO and the case-studies.

Postmark. For an I/O-intensive workload, we used Postmark [VERITAS Software 1999], a popular file system benchmarking tool. Postmark stresses the file system by performing a series of file system operations such as directory lookups, creations, and deletions on small files.

TPC-C. TPC-C [Transaction Processing Performance Council 2004] is an On-Line Transaction Processing (OLTP) benchmark that performs small 4 KB random reads and writes. Two-thirds of the I/Os are reads. We set up TPC-C with 50 warehouses and 20 clients. We compare our context-aware MySql running on our CAIO framework with regular MySql running on a vanilla kernel. The metric for evaluating TPC-C performance is the number of transactions completed per minute (tpmC). We report tpmC numbers for each benchmark.

Results. Figure 14 shows the overheads of our CAIO framework for Postmark for two different number of operations. As seen from the figure the overall elapsed time overheads were small (2% to 4%) compared to regular I/O. This overhead is mainly because of the additional user-to-kernel copies for communicating context objects from applications.

TPC-C Results. The TPC benchmark results for regular MySQL and our modified context-aware MySQL ran over the CAIO kernel is shown in Table I. The workload loads tables into a Mysql server at start-up and runs a mix of queries on these tables for a user defined time. As seen from throughput and response time numbers, overheads of the CAIO framework is quite small.



(a) 50,000 Operations (b) 100,000 Operations

Fig. 14. *Postmark Results for CAIO Framework*

	Regular Response Time (s)	**CAIO Response Time (s)**
Delivery	0.096	0.109
New Order	0.039	0.064
Order Status	0.033	0.29
Payment	0.000	0.000
Stock Level	0.169	0.524
Throughput (tpmC)	67.13	64.35

Table I. *TPC-C Benchmark results for the CAIO framework*

7. CASE STUDY: WORKING SET IDENTIFIER

Our first case study is the automatic **Work**ing **S**et **IDE**ntifier (*WorkSIDE*). WorkSIDE that uses both access-bound and data-bound contexts to automatically infer the minimum set of data items required to be available in order for an application (or a specific instance of an application) to run to completion.

7.1 Motivation

This ability to accurately identify working sets of application contexts at a fine grained level has various kinds of applications.

Performance. The working set of the application can be preloaded into a much faster but smaller memory hierarchy (e.g., a flash storage layer that provides about 100x better random access read performance), thus essentially shielding the application from performance variability due to disk access.

Availability. WorkSIDE enables fault-isolated placement of application working sets enabling truly graceful degradation during multiple disk failures similar to D-GRAID [Sivathanu et al. 2004]. While D-GRAID could just co-locate files or directories, WorkSIDE can co-locate higher-level application working-sets within failure domains.

Power Savings. Many recent systems have looked at saving power by switching off a subset of disks in a large RAID array in such a way that applications can still function

properly without the switched-off disks [Zhu et al. 2005; Weddle et al. 2007]. These systems go to great complexity to identify the subset of data that is currently under use, yet these techniques are most often approximate and too coarse-grained. Being more informed about the application's access patterns and data abstractions, WorkSIDE can do a better job at such power optimizations by being more aggressive and more accurate.

Disconnected operation. Another usage scenario for WorkSIDE is when the user wants to preload the working set for a specific application context in local storage for disconnected operation, say, in a mobile environment. This enables Coda-like hoarding [Kistler and Satyanarayanan 1991], but can be much more accurate, fine-grained and automated. For example, if the user works only on a specific build target in a large body of source code, just the subset of source files (and the metadata) needed for the target can be automatically preloaded to local storage.

The key to WorkSIDE is its ability to correlate a repeatable access context with the data context it accesses. WorkSIDE achieves this by associating with each node of the access context hierarchy, the aggregated set of data items that are accessed by that context. Semantic aggregation of such data is possible because data-bound contexts are hierarchical in nature conveying data abstractions in several granularities (such as files or directories). Tracking working set at an aggregated level enables much simpler and reliable tracking of repeatability. For instance, if an application touches different parts of a file in its different runs, block-level tracking may not find much of a repeatability, whereas tracking at the file-level would indicate the pattern. Since the data context hierarchy essentially contains information of the entire data abstraction tree, it can track this information at various granularities, and decide on which granularity provides the best trade-off between the amount of data to be preloaded and ensuring completeness for the application.

7.2 Design

To determine the working set of a higher level logical task, WorkSIDE has to track history of both data-bound and access-bound contexts for every task. We designed WorkSIDE as an on-disk mechanism to demonstrate its working as part of the firmware of a high-end block-based RAID storage system. WorkSIDE can potentially exist at any layer of the storage stack such as the file system or the device driver. Through our design, we show that even in the lowest layer of the storage stack (the storage hardware), working set identification can be done to an acceptable level of accuracy, through context-aware I/O.

For WorkSIDE to correctly determine the working set of data for a given access-bound context, the higher application has to pass data contexts to communicate the semantic organization of data. This can relate to on-disk structures such as B-trees, database tables, files, and directories. In this section, we first detail how access-bound contexts can be associated with corresponding data-bound contexts. We then discuss a few policies that can be adopted to determine the granularity of the working set of a given context. Lastly, we present our prototype implementation of WorkSIDE.

Associating Access with Data. WorkSIDE maintains two context stores (described in Section 6.4) to track access-bound and data-bound contexts respectively. Each store has context trees to represent the hierarchy. We call tree nodes in the access and data stores as *Access-Context Nodes* (ACNs) and *Data Context Nodes* (DCNs) respectively. Note that, as data-bound context is mainly used to communicate the semantic structure of data, it need not necessarily be passed by the higher-level application for every I/O request. For

example, if a DBMS uses the `table` and `record` abstractions as data-bound contexts, it may pass the context hierarchy only when such abstractions are created (e.g., a table creation) or updated (e.g., a new record insertion). For example, the DBMS need not pass data-bound contexts for every `select` query. To handle this condition, WorkSIDE may have to map access-bound contexts accompanying a block I/O request with a pre-existing data-bound context hierarchy.

The following are the contents of a DCN: (a) A context identifier. (b) The number of blocks in the entire sub-tree with the node as root. (c) A list of block numbers associated with the context (if it is a leaf node). Every time a block I/O has an accompanying data-bound context chain, the corresponding block number is added to the leaf DCN of the chain. (d) A list of pointers to its child nodes. (e) A back-pointer to its parent node. This is used to increment the number of blocks in every parent along the chain when there is a new addition to a leaf node.

While adding a node to the tree, we enforce the *single parent* constraint, where every node must have at most one parent. When there is a context chain passed, that violates this condition, we truncate the chain after the spurious node while adding it to the tree. In almost all common cases, this would not affect the accuracy of the data-bound context tree, as most data-abstractions already follow this rule. For example, a single block cannot belong to more than one file (except in rare cases such as hardlinks in Ext2).

WorkSIDE also maintains a hash table, BDTABLE, to map block numbers to the corresponding leaf nodes in the data context tree. The BDTABLE is used to lookup the data context for any block when an I/O request to it does not have an associated data-bound context. Upon receiving a block I/O request with a access-bound context, WorkSIDE can map the corresponding block number to any level of abstraction in the data-bound hierarchy by just traversing through the parent back-pointers in each node in the data context tree.

In the next section, we describe how this infrastructure is augmented with association policies to determine the optimal granularity of associating a data-bound working set for a given access-bound context.

Working Set Identification. Identifying the working set for a given node in the access-bound context tree involves associating that ACN with one or more DCNs. Therefore every ACN in the access store contains pointers to one or more DCNs.

Greatest-Common-Prefix Mode. We designed WorkSIDE to operate under two different modes for choosing the appropriate DCN for a given ACN. In the first (and simple) mode, which we call the *Greatest Common Prefix* (GCP) mode, WorkSIDE maintains utmost one DCN per ACN. Whenever there is an I/O in the context of an ACN, the request block number is looked up in the BDTABLE to find the leaf DCN to which the block number is associated. The leaf DCN is associated with the ACN if the ACN did not previously have a DCN associated. If not, the greatest common prefix node in the tree (starting from the root) for the new leaf DCN and the previously associated DCN is computed (using the parent back-pointers) and associated with the ACN. The working-set is enumerated by just traversing the sub-tree starting from the associated DCN. This method of enumerating the working set for an ACN ensures completeness, but under some scenarios there could be a significant number of falsely associated blocks. For example, if an access context A reads files `/home/john/plan.txt` and `/home/john/private/list.txt`, the

GCP method of association would include the entire contents of `/home/johh/` in the working set of A. A variant of the GCP mode mitigates this problem under some scenarios by tracking the longest depth to traverse while enumerating blocks, along with the ACN. With this, the working set of A would just include files up to depth level 3 (`/home/john/private`).

Multi-DCN Mode. In the second mode, which we call the *Multi-DCN mode*, WorkSIDE tracks a list of DCNs per ACN. Every ACN has a list of duplicate eliminated pointers to parent DCNs. To enumerate the working set for a given ACN, the following procedure is used: for each DCN associated, all blocks belonging to their immediate children are included. For example, if an ACN B reads files `/home/john/plan.txt` and `/home/john/private/list.txt`, DCNs for `/home/john` and `/home/john/private` will be associated with B. While enumerating the working set of B, all files (not subdirectories) under `/home/john` and `/home/john/private` will be included. Therefore, the multi-DCN mode of association provides more accurate identification of working sets. However, this method needs to track more information per ACN. In the procedure described above, we choose the hierarchy one level above the leaf DCN for every block access. However, the number of such levels can be configurable based on specific system and workload requirements.

WorkSIDE can also track information required for both GCP and multi-DCN modes simultaneously (every ACN can have both the list of parent DCNs and a single GCP node). Based on the kind of usage scenario for the working set, enumeration process can be decided dynamically to choose the optimal granularity.

Prefetcher. We developed an on-disk prefetching tool that uses WorkSIDE to enumerate the working set of access-bound contexts and prefetch them into a faster storage. For prefetching, we tracked the repeatability of the working set of each ACN, and for repeatable ACNs, we prefetch and serve the entire working set from the faster storage medium. Currently we use RAM to cache prefetched working sets, but this could even be a fast secondary storage device such as flash. While deciding whether to prefetch a working set, we take into consideration the size of the working set and the available space in the prefetch cache. In our design, we use a simple scheme where we prefetch working sets less than half the size of the prefetch cache subject to remaining space availability in the cache. More advanced algorithms such as best-fit and worst-fit can also be implemented to decide the appropriate working sets to prefetch.

To evaluate our working-set aware prefetcher, we compiled several modules in the Linux kernel source, and `e2fsprogs` package [Ts'o 2008], with inheritable contexts. We found that once working-sets were identified by WorkSIDE and prefetched into RAM by our prefetcher, there were no requests sent to the disk during the compile workload. Therefore, working-set aware prefetching of data enables turning off disk drives (and hence save power) in the case of repeatable workloads.

7.3 Implementation

We implemented a prototype of WorkSIDE and our prefetching tool as a pseudo-device driver in Linux kernel 2.6.15 that stacks on top of an existing disk block driver. The pseudo-device driver receives all block requests, and redirects the common read and write requests to the lower level device driver, after storing context information that needs to be tracked. Our prototype of WorkSIDE included both the GCP and multi-DCN modes

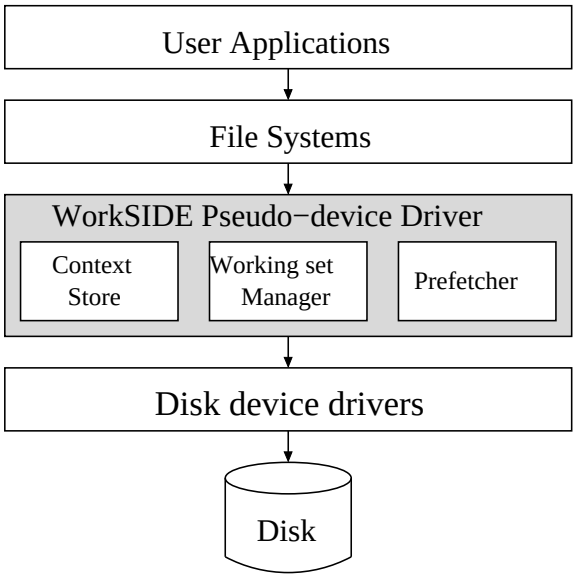


Fig. 15. *Prototype implementation of WorkSIDE in the Linux kernel. The prefetcher component in WorkSIDE prefetches common working-sets into memory to save power.*

of associating data-bound contexts. It contains 3,020 lines of new kernel code. Figure 15 shows the architecture of our prototype.

For testing WorkSIDE, we also modified the VFS layer of the Linux kernel to encode the pathname of the entity being operated (file or directory) along with every lower level I/O request. File system meta-data blocks such as super blocks, bitmaps and directory blocks have to be dealt with separately, as they may not particularly belong to a specific application. To handle such blocks, we modified the Linux Ext2 file system to associate a generic "common" context which can be interpreted by any layer as one that is not associated with any particular access-bound context. We call our modified Ext2 file system, Ext2C.

7.4 Evaluation

We evaluated the correctness and performance of our prototype implementation of WorkSIDE. For correctness we used a Linux kernel module build process, and for performance, we used the Postmark benchmark described under Section 6.4.

Completeness of the Working-Set. To verify the completeness of the working-sets identified by WorkSIDE, we implemented a prefetch cache layer beneath the file system that prefetches the working-set for selected access-bound contexts. We then simulated a disk crash by disallowing disk I/O from our pseudo-device driver, and repeated the workloads for the corresponding contexts. We performed this for kernel modules compiles and several micro-benchmarks, and in all cases the prefetch cache layer serviced all I/O requests. This shows that working-sets identified by WorkSIDE are complete.

Kernel Modules Build. Our goal during this test was to evaluate the correctness of the working set identification mechanism of WorkSIDE. We untarred a vanilla Linux 2.6.15 kernel on our Ext2C file system mounted over our WorkSIDE pseudo-device driver. We did

Module	# Directories	# Files	# Blocks (4k)
Ext2	14	315	1149
Ext3	14	328	1452
ReiserFS	14	328	1432
NTFS	14	320	1769

Table II. Compilation Working Set Statistics

this through a shell that has an inheritable access-bound context set (described under Section 6.4), with depth one. We then remounted the file system to eliminate cache effects and compiled the source-code of a few file systems (Ext2, Ext3, Reiserfs, and Ntfs) under the `fs/` sub-directory of the kernel source. While compiling each file system, we used different shells with different second-level inheritable contexts set. All compilations were done with the same top-level hierarchy of context, but for each compilation, the second-level was different. Therefore, we were able to track the working-set of each of the individual compilations. Note that we initialized the build process through "`make install`" separately at the beginning, and remounted the file system after each compilation. We ran this test over WorkSIDE for both GCP and multi-DCN modes of operation.

Under the GCP mode, we noticed that the working sets of every single file system compilation was identified as a the root of the kernel source tree. This is because, a file system module compilation would refer to files under `include/` and `fs/` and hence the greatest common prefix node becomes the root of the kernel source.

When we ran the test under the multi-DCN mode, we saw WorkSIDE identify separate working sets for each of the file system compilation contexts. Table II shows the total number of directories, files, and blocks associated with the working set of each compilation. We identified these by dumping the entire access-bound context tree of WorkSIDE and their associated DCNs. In each compilation context, the generated object files were also included in the working set as the same inheritable context was passed for write operations as well.

We also used the Multi-DCN mode of WorkSIDE to calculate the working-sets for kernel compilation with `make allnoconfig` and `make allyesconfig`. For compilation using `make allnoconfig`, the size of the working-set came out to 32.6MB. For `make allyesconfig`, the working-set size was 3GB. As the object files during compilation are created from the same context, they were included in the working-set.

Postmark. To evaluate the performance overheads of WorkSIDE, we used an I/O-intensive benchmark, Postmark. We ran our modified Postmark that passes context objects with each I/O request, over WorkSIDE in its two modes, and compared it with regular Postmark running on top of a normal disk. For the regular Postmark we used unmodified Ext2 as the file system and for WorkSIDE evaluation, we used our modified Ext2C file system. Figure 16 shows the overheads of WorkSIDE compared to regular disks.

WorkSIDE under the GCP mode of operation had an elapsed time overhead of 1.5% compared to regular disk. The overhead mainly consists of system time (12%) caused because of updating context trees and tracking greatest common prefixes. Under the multi-DCN mode of operation the elapsed time overhead was 3.7% compared to a regular disk, caused by a 20% increase in system time. The increase in overheads compared to GCP mode is because under the multi-DCN mode, WorkSIDE has to track multiple data nodes per access-node. If WorkSIDE is implemented in a real disk, tracking context trees would

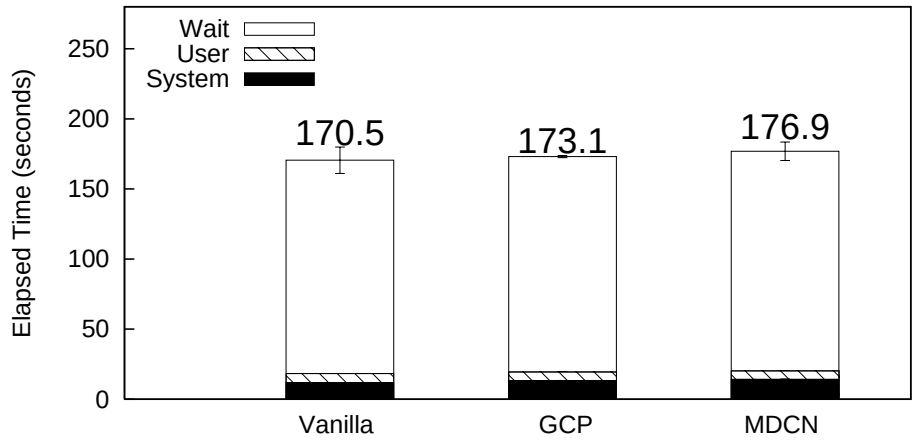


Fig. 16. *Postmark Results for WorkSIDE (200 Sub-directories, 20,000 Files, and 200,000 Transactions.). This shows the overheads associated with the process of working-set identification at the disk-level.*

be done by the disk firmware and hence would not incur the host CPU overheads.

8. CASE STUDY: CONTEXT-AWARE CACHING

Modern large-scale storage systems have hundreds of gigabytes of built-in main memory [EMC Corporation 1999], primarily for caching purposes. However, today's storage systems cannot adapt their caching policies based on application-level workloads or data semantics, as they lack information about higher level semantics. This is caused by an excessively simple disk interface [Denehy et al. 2002; Sivathanu et al. 2004]. Application-aware caching policies have been found to be quite useful in the context of OS level caches [Cao et al. 1996]. Yet today's disk systems cannot even separate independent I/O streams generated by two different applications, making it harder to implement application-aware caching policies.

In this section we design and evaluate *Context-Aware Cache* (CA-*cache*), an on-disk caching mechanism that differentiates independent I/O streams using logical contexts and tunes its caching policies based on individual access patterns.

8.1 Design

We designed CA-*cache* as an on-disk LRU write-through cache layer. The goal of CA-*cache* is to identify sequential streams of I/O and disable caching their data, as mostly sequential I/O streams do not benefit from read caching. As we are interested in the access-patterns to tune the caching policy, this application uses access-bound contexts.

Architecture. CA-*cache* consists of a set of dynamically-built context trees and an LRU cache. Each tree represents a group of hierarchical contexts with the same root context. Each node represents the hierarchical context specified by the path from the root of the tree to that node. Context trees are created or updated on each read request that specifies an access-bound context.

Classification of Contexts. Each node in the tree contains the following information about a particular context: (a) the inferred access-pattern for the particular context, (b) the block number for the last read I/O request required to track sequentiality, and (c) two counters that track the number of successive sequential and random read requests in the past. A context node is initialized as random-access upon creation. Based on the last read request and the current request, either the sequential or the random counter is incremented and the other is reset. When the values of the counters exceed a *threshold*, the node is classified as sequential or random as appropriate. Note that an already classified node could be re-classified when its access pattern changes. Upon receiving any read request, the counters in all nodes that are part of the current context are updated and the nodes are re-classified if needed. We call the number of sequential read request required for classifying a node as sequential, the *sequential threshold*. The sequential threshold is configurable, and can range somewhere between 10 and 100. A sequential-access node is re-classified as random upon a single out-of-order read.

Caching Methodology. Our classification scheme allows for different hierarchy levels in the same context chain to be classified differently. For example, two sub-contexts that are part of the same parent may be doing sequential I/O in their own levels. However, since the I/O from the sub-contexts could be received interleaved, the parent would be classified as random. CA-*cache* does not require context identifiers to be repeatable. Therefore, it contains a mechanism to automatically forget contexts based on a timeout. We periodically purge context tree entries that represent inactive contexts (without any requests) beyond a time threshold.

8.2 Evaluation

We implemented a prototype of our on-disk caching mechanism as a pseudo-device driver in the Linux 2.6.15 kernel similar to WorkSIDE. We maintain the context trees in memory and an asynchronous kernel thread wakes up periodically to purge timed out context entries. If the block is present in the LRU cache, the pseudo-device driver services the request from the cache, thereby avoiding a request to the lower level. Otherwise, the request is directed to the lower level and the cache is updated on completion of the request, if the request belongs to a random-access context.

Read Micro-benchmark. To evaluate CA-*cache*, we ran a micro-benchmark that generates synthetic random and sequential read workloads simultaneously and calculated the overall throughput of the random workload. We compared the throughput results of CA-*cache* with a vanilla LRU cache layer which treats all contexts equally. Both CA-*cache* and vanilla LRU cache used 4MB of cache (1,024 4KB pages) for this benchmark.

We ran a user program that generates workloads shown in Figure 17. The user program has four execution contexts (threads), A, B, C, and D which use their own files for I/O. Thread A reads a 4GB file sequentially with context {1–2–5} (see Figure 17). Thread B reads a 4GB file sequentially, but it uses contexts {1-3-7} and {1-3-8} for alternate reads. Thread C is identical to thread B, but it uses contexts {1–4–9} and {1–4–10}. Thread D reads random locations from a 4GB file using context {1–2–6}. For thread D, we use a random number generator that repeats itself every 1,024 reads. The threads run until any one of the sequential threads exits after reading 4GB of data. In our experiment, the throughput of the random workload when run under the vanilla LRU cache was 0.098 MB per second, whereas with CA-*cache*, the throughput was 7.71 MB/Sec.

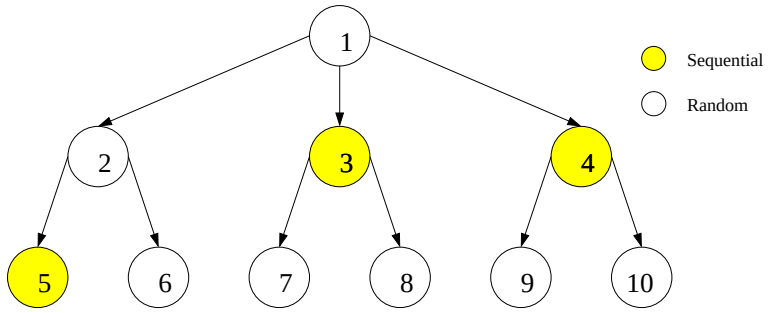


Fig. 17. *Context tree used for CA-cache micro-benchmark. After our micro-benchmark, CA-cache classified the grayed nodes as sequential and the rest as random.*

MySQL Micro-benchmark. For this benchmark, We created two identical tables SEQ and RAND in MySQL with 4,200,000 records each, and ran random and sequential query logs simultaneously. The tables were approximately 233MB in size. The sequential query log contained a `select *` query on the table. For a random workload, we selected a subset of the records at random and issued select queries based on their record IDs. To show the benefits of caching random streams alone, we repeated the random query log ten times. We also ran the sequential log in a loop till the random workload completed. We determined the throughput of the random workload (number of queries executed per second) while the sequential workload was running in parallel. It was 266.13 queries per second without selective caching, while it was 614.15 queries per second with selective caching.

9. RELATED WORK

In this section, we discuss related research for the concepts, techniques, and insights used in our abstractions and the case-studies that we developed.

9.1 Briding the Information-Gap in the Storage Stack

Several systems have been proposed with the overall goal of bridging the information-gap in the system stack. In this section, we classify existing research in this area into four categories: extensible systems, richer abstractions, hint-based interfaces, and inference-based systems. The related work for the case-studies for each abstraction is discussed under their respective sections.

Extensible Systems. Building extensible systems are a solution to the problem of information-gap in the storage stack. Extensible operating systems [Bershad et al. 1995; Seltzer et al. 1994] allow applications to implement their own policies for traditional operating system tasks, by ensuring a safe execution environment for them. A related approach is the one taken by Exokernel [Engler et al. 1995], which advocates building a minimal operating system and have everything else be implemented in application libraries.

The notion of extensibility has also been explored at the hardware level. For example, active disks [Acharya et al. 1998; Riedel 1999] enable applications to download code into the disk that is run within the disk controller. Such code can implement arbitrary filtering of data based on application level predicates, and even perform more sophisticated operations such as search [L. Huston and R. Sukthankar and R. Wickremesinghe and M.

Satyanarayanan and G. R. Ganger and E. Riedel and A. Ailamaki 2004] without actually transferring data out of the disk subsystem. Scriptable RPC [Sivathanu et al. 2002] proposes making the interface of a network file server extensible so that clients can dynamically implement flexible cache consistency and concurrency policies.

All these systems provide a lot of control to the application and in the process, essentially ties them together. For applications to actually use such extensible layers, they need to have a reasonably intricate understanding of the system, thus making them complex to design. Nevertheless, for applications that really require such control and can utilize it sensibly, these provide the right level of abstraction.

Hint-Based Interfaces. A more evolutionary approach that past research has explored is to provide specific primitives at the system level that the applications can use to convey information to the operating system. Informed prefetching [Tomkins et al. 1997] is an example of such a system. By enabling the application to convey information on its future access pattern, the operating system acquires knowledge about the application that it uses to perform more intelligent prefetching. Another example is the Logical disk [de Jonge et al. 2003], which provides an interface for the applications to encode locality hints by creating lists of blocks. Researchers have also looked at the flip-side of the problem: provide information about the operating system to the application so that the application can make intelligent decisions. Infokernel [Arpaci-Dusseau et al. 2003], and icTCP [Gunawi et al. 2004] advocate the approach of the operating system exporting a minimal amount of internal information which the applications then use to tune their behavior.

Previous work has also looked at the idea of conveying application knowledge through new abstractions. Perhaps the closest to our work is the idea of Resource Containers [Banga et al. 1999], which allows applications to group requests into a resource container which is then treated as a logical principal for the purposes of resource isolation. However, even Resource Containers were built with the specific goal of resource accounting.

One commonality between many of these hint-based approaches is that the hints are often tied to a specific kind of optimization or functionality. In other words, the information being transferred is designed with a particular purpose in mind. This in turn limits the flexibility of such a system because each new class of functionality may require yet another new primitive to be added to the interface.

Richer Abstractions. Our work is closely related to a large body of work examining new interfaces between file systems and disk storage. For example, logical disks expand the block-based interface by exposing a list-based mechanism that file systems use to convey grouping between blocks [de Jonge et al. 2003]. The Universal File Server [Birrell and Needham 1980] has two layers where the lower layer exists in the storage level, thereby conveying directory-file relationships to the storage layer. More recent research has suggested the evolution of the storage interface from the current block-based form to a higher-level abstraction. Object-based Storage Device (OSD) is one example [Mesnier et al. 2003]; in OSDs the disk manages variable-sized objects instead of blocks. Object-based disks handle block allocation within an object, but still do not have information on the relationships across objects. Another example is Boxwood [MacCormick et al. 2004]; Boxwood considers making distributed file systems easier to develop by providing a distributed storage layer that exports higher-level data structures such as B-Trees. ExRAID [Denehy et al. 2002] explores the utility of exposing hardware specific informa-

tion from a RAID device to the higher layers such as the file system.

These interfaces are designed with some specific applications or scenarios in mind. For example, it is hard to implement a database in an object-based disk. This illustrates that it is hard to design a generic interface that is suitable for a wide-range of applications.

Inference-Based Systems. Inference-based systems take the extreme approach of making no modifications to interfaces, but *infer* cross-layer information without explicit transfer for information across the layers.. Gray-box systems [Arpaci-Dusseau and Arpaci-Dusseau 2001] is an early example of such an approach. An application with "gray-box" knowledge of the operating system attempt to implicitly control the operating system behavior by tuning its workload in such a way that it takes the operating system to a state that results in the desired policy. Another system built along the same philosophy is semantically-smart disks [Sivathanu et al. 2004] in which the storage system infers knowledge about the higher layers by carefully observing traffic patterns and correlating them to higher level operations.

Although inference-based techniques are valuable from the viewpoint of being easily deployable and less intrusive, these approaches have their own limitations because they are heavily constrained in terms of not changing interfaces. This in many cases results in additional complexity, making it hard to reason about correctness while also limiting the usage of such inferred knowledge to less aggressive applications that can tolerate inaccuracy.

9.2 Type-Safety

The concept of type safety has been widely used in the context of programming languages. Type-safe languages such as Java are known to make programming easier by providing automatic memory management. More importantly, they improve security by restricting memory access to legal data structures. Type-safe languages use a philosophy very similar to our model: a capability to an encompassing data structure implies a capability to all entities enclosed within it. Type-safety has also been explored in the context of building secure operating systems. For example, the SPIN operating system [Bershad et al. 1995] enabled safe kernel-level extensions by constraining them to be written in Modula-3, a type-safe language. Since the extension can only access objects it has explicit access to, it cannot change arbitrary kernel state. More recently, the Singularity operating system [Hunt et al. 2005] used a similar approach, attempting to improve OS robustness and reliability by using type-safe languages and clearly defined interfaces.

9.3 Capability-Based Access Control

Network-Attached Secure Disks (NASDs) incorporate capability based access control in the context of distributed authentication using object-based storage [Gibson et al. 1998; Aguilera et al. 2003; Miller et al. 2002]. Temporal timeouts in ACCESS are related to caching capabilities during a time interval in OSDs [Azagury et al. 2003]. The notion of using a single capability to access a group of blocks has been explored in previous research [Gobioff 1999; Miller et al. 2002; Aguilera et al. 2003].

In contrast to their object-level capability enforcement, ACCESS uses implicit path-based capabilities using pointer relationships between blocks.

9.4 The Notion of Context in Storage

The idea of tagging requests with identifiers has been explored in the context of distributed systems for performance debugging, profiling, etc. Pinpoint [Chen et al. 2002] and Magpie [Barham et al. 2003] are examples of systems in this category. Recently, Thereska et al. proposed applying a similar idea in the context of distributed storage systems mainly for performance monitoring [Thereska et al. 2006]. All these systems look at tagging requests in a causal chain with a certain identifier so that the entire *path* of a logical request (which may involve multiple physical network hops) can be tracked. Researchers have also looked at implicitly inferring this causal knowledge without explicit tagging [Aguilera et al. 2003; Gniady et al. 2004; Li et al. 2004] but it involves significant complexity compared to the explicit tagging approach. These systems only operate within the scope of one logical request and are targeted at a specific application. In contrast, CAIO allows for a more general expression of application level semantics to cater to a wide variety of applications.

Previous work has also looked at conveying application-level grouping through new abstractions similar to our notion of context. Perhaps the closest to our work is the idea of Resource Containers [Banga et al. 1999], which allows applications to group requests into a resource container which is then treated as a logical principal for the purposes of resource isolation and accounting. However, similar to the systems discussed above, resource containers were also built with the specific goal of resource accounting and convey information on one specific kind of grouping.

Our work on context-aware I/O also fits into a class of other work on general solutions for bridging the information gap across system layers. Work in this area mainly belongs in three categories: extensible systems, hint-based interfaces, and implicit techniques to infer information or exert control.

9.5 File System Consistency

Consistency mechanisms for file systems have been explored extensively. Early file systems such as FFS [McKusick et al. 1984] relied on a global scan of disk metadata to fix consistency problems. This mechanism, called the file system consistency check (fsck) was in popular use until recently in the Linux Ext2 and Windows VFAT file systems. However, as increasing disk sizes made such global scans more and more expensive, more efficient mechanisms have become popular. Journalling, originally proposed as early as in the Cedar file system [Gifford et al. 1988], uses database like transactions for metadata updates. Modern file systems such as Ext3 and Windows NTFS use journalling for file system consistency. Another technique proposed for file system consistency is Soft Updates [Ganger et al. 2000; McKusick and Ganger 1999], which orders updates carefully so that pointer dependencies get updated in the right order. Soft updates is somewhat similar in spirit to our approach since it is also pointer-based. A relatively recent study evaluated the trade-offs between journalling and soft updates [Seltzer et al. 2000].

Database systems have for long used mechanisms for consistency. Consistency in databases is enforced via transactions; the ARIES transaction based recovery mechanism [Mohan et al. 1992] is used quite widely in database systems. The basic technique is to group all related updates into a single transaction that is then committed to disk atomically, so that the state remains consistent. As we described in Section 5.6, transactions are more general and powerful than pointer-based consistency, but using transactions requires a fair bit of work at the application level. Our mechanism provides a simpler yet effective alternative

to transactions, although not as general.

Consistency at the disk level has been explored in the context of Semantically-smart disks (SDS) [Sivathanu et al. 2003]. In that paper, the authors implement journalling underneath unmodified Ext2 by utilizing inferred semantic knowledge. However, in their work, the disk system had to be aware of the specific structures at the file system level and thus was tied to a specific file system. Further, it required a synchronous mount of the file system. Our work explores enforcing consistency in a manner generic to the higher level software. However, in the process, we require changing the file system or software above to use the pointer API. We therefore view both these approaches as complementary.

10. CONCLUSIONS

As Butler Lampson said, interface design is one the most complex aspects of system design, while also being the most important. Interface designers have traditionally embraced the philosophy of minimalism—hide as much information about the layers as possible, so that the layers can innovate and evolve independently. This approach, despite all its merits, has the downside of obscuring what a layer knows about its inputs, thus limiting functionality. At the other extreme, some systems have explored how to completely tie the layers together, by having extensible layers, or exposing detailed information about the inner semantics of a layer. What we have explored in this article is a middle-ground, where we send a small amount of information across layers. By making the generation of the information separate from how the information is used, we enable the layers to be independent of each other, while still enabling arbitrary grouping and relationships to be conveyed across the storage stack.

10.1 Lessons Learned

We now discuss four key lessons learnt through our experience in evolving and prototyping our end-to-end abstractions and the case-studies. We believe these lessons would be useful for future interface designers not only in the storage domain, but also more generally in computer systems.

> ***Lesson 1: Generalizing structural and operational information in storage is possible.***

Our pointer abstraction shows that higher-level structures such as files, directories, database tables, or B-trees can be formalized in a generic manner by way of pointers. The fundamental insight behind the pointer abstraction is that today's disk systems store data in the form of fixed size blocks. Therefore, to implement higher-level structures on top of this simple abstraction, relationships have to be established between these individual blocks. Most file systems and other storage software today maintain these relationships through *explicit* pointers. Even if pointers are *implicit* as in the case of extent-based storage design, it is straightforward to generate them explicitly for communicating to the storage stack.

The context-aware storage abstraction provides a means to formalize *operational* information in addition to structural knowledge. By way of hierarchical context identifiers, we show how application-level operational contexts can be encoded in a generic manner even for complex storage applications such as databases.

> ***Lesson 2: Requiring just implementation-level modifications to existing infrastructures is a virtue in interface design.***

Both our abstractions require only implementation-level modifications to existing software layers. Our straightforward implementations of the Ext2TSD and VFATTSD file systems that support the type-aware storage abstractions indicate that as long as there is not a need to redesign existing infrastructures, interface changes are easy to be adopted and deployed. The limited changes that we made to the MySQL and the Linux kernel to support hierarchical contexts corroborate this fact.

Lesson 3: Annotating pointers or contexts with application-level attributes enables a wider range of functionality.

To support new features that need to be tuned for specific applications or storage layers, annotating generic information with optional *attributes* proves to be useful. These attributes need not be part of the main interface, but can be communicated *offline* between specific layers. For example, the share proportion for different contexts in our a proportional-share disk scheduler can be set offline by the administrator, specifically in the disk scheduler layer.

Lesson 4: Decoupling the generation of information from its usage has its own limitations

Although our abstractions enable a wide-range of new functionality in the storage stack, they cannot support certain kinds of features that require precise application-specific information. For example, although type-awareness enables disks to group blocks based on pointers, disks cannot precisely identify if a particular group represents a file, directory, or a database table. Although it is true that a large class of new functionality can be achieved without such knowledge, some features that needs to use more fine-grained application-specific information cannot be implemented without help from applications. Similarly, although context-aware storage encodes all granularities of application contexts, lower layers cannot identify what each level in the hierarchy means, which may be needed for certain functionality.

10.2 Future Work

In this section, we discuss potential future directions to explore in the topic addressed by this article. We first talk about how the general principle behind our abstractions can extend more broadly in other domains. We then discuss two possible future directions to develop new applications using our abstractions.

Generalizing Information in Other Domains. What we have explored in this article is how structural and operational knowledge about application data can be formalized and used to bridge the information-gap in the storage stack. This general principle of formalizing the different *properties* of application data is relevant in other domains. For example, it could be interesting to explore if security polices can be formalized in a minimal and generic manner and propagated across the systems stack to enable a new class of secure systems, where different layers can independently provide security features without explicit coordination from applications.

Applications in Virtual Machine Environments. The growing popularity of virtual machine technology exacerbates the problem of information-gap in the systems stack, as it introduces another layer of virtualization. Bridging the gap in this context enables highly useful optimizations and new functionality at the virtual machine host. For example, if the

host kernel is aware of structural information about data in virtual machines, it can implement security features such as global anti-virus checking, intrusion detection, or access control, that cannot be bypassed by any guest virtual machine.

Applications in Distributed Environments. Distributed environments present a similar scenario as virtual machines in the aspect of information-gap. We believe that the notion of hierarchical contexts enables a wide range of functionality in distributed systems. Starting from straightforward features such as distributed performance isolation, contexts can potentially go a long way in enabling more complex and interesting functionality such as custom reliability and consistency policies and so on.

10.3 Summary

Overall, we find that type-awareness and context-awareness in the storage stack enables an interesting set of new functionality and optimizations, with minimal modifications to existing infrastructures. We believe that our abstractions explore an interesting and effective design choice in the large spectrum of work on alternative interfaces to storage. As described in Section 10.2, we believe that the insights derived in this article apply broadly in several other systems domains.

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