Barrier Enabled IO Stack for Flash Storage

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Abstract

This work is dedicated to eliminating the overhead of guaranteeing the storage order in modern IO stack. The existing block device adopts prohibitively expensive re-sort in ensuring the storage order among write requests: interleaving successive write requests with transfer and flush. Exploiting the cache barrier command for the Flash storage, we overhaul the IO scheduler, the dispatch module and the filesystem so that these layers are orchestrated to preserve the ordering condition imposed by the application till they reach the storage surface. Key ingredients of Barrier Enabled IO stack are Epoch based IO scheduling, Order Preserving Dispatch, and Dual Mode Journaling. Barrier enabled IO stack successfully eliminates the root cause of excessive overhead in enforcing the storage order. Dual Mode Journaling in BarrierFS dedicates the separate threads to effectively decouple the control plane and data plane of the journal commit. We implement Barrier Enabled IO Stack in server as well as in mobile platform. SQLite performance increases by 270% and 75%, in server and in smartphone, respectively. Relaxing the durability of a transaction, SQLite performance and MySQL performance increases as much as by $73 \times$ and by $43 \times$, respectively, in server storage.

1 Motivation

Modern IO stack is a collection of arbitration layers; IO scheduler, command queue manager, and storage write-back cache manager. Despite the compound uncertainties from the multiple layers of arbitration, it is essential for the software writers to ensure the order in which the data blocks are reflected to the storage surface, storage order, e.g. in guaranteeing the durability and the atomicity of a database transaction [46, 26, 35], in filesystem journaling [63, 40, 64, 4], in soft-update [41, 61], or in copy-on-write or log-structure filesystems [59, 35, 58, 31]. Preserving the ordering requirement across the layers of the arbitration is being achieved by an extremely expensive resort; dispatching the following request only after the data block associated with the preceding request is completely transferred and is made durable. We call this transfer-and-flush mechanism. For decades, interleaving the writes with transfer-and-flush has been the fundamental principle to guarantee a storage order in a set of requests [23, 15].

The concurrency and the parallelism in the Flash storage, e.g. multi-channel/way controller [70, 6], large size storage cache [47], and deep command queue [18, 27, 69] have brought phenomenal performance improvement. State of the art NVMe SSD reportedly exhibits up to 750 KIOPS random read performance [69], which is nearly $4,000 \times$ of HDD’s performance. On the other hand, the time to program a Flash cell has barely improved if it has not deteriorated [21]. This is due to the adoption of the finer process (sub 10 nm) [24, 36], the multi-bits per cell (MLC, TLC, and QLC) [5, 10] in the endless quest for higher storage density [42]. Despite the splendid performance improvement of the Flash storage claimed by the storage vendors, the service providers have difficulty in fully utilizing the underlying high performance storage.

Fig. 1 alarms us an important trend. We examine the
performance of write with ordering guarantee (write() followed by fdatasync()) against the one without ordering guarantee (write()). We test seven Flash storages with different degrees of parallelism. In a single channel mobile storage for smartphone (SSD A), the performance of ordered write is 20% of that of the buffered write. In a thirty-two channel Flash array (SSD G), this ratio decreases to 1%. In SSD with supercap (SSD E), the ordered write performance is 25% of that of the buffered write. There are two important observations. First, the overhead of transfer-and-flush becomes severe as the the degree of parallelism increases. Second, use of Power-Loss Protection (PLP) hardware fail to eliminate the transfer-and-flush overhead. The overhead is going to get worse as the Flash storage employs higher degree of parallelism and denser Flash device.

Fair amount of works have been dedicated to address the overhead of storage order guarantee. The techniques deployed in the production platforms include non-volatile writeback cache at the Flash storage [22], no-barrier mount option at the EXT4 filesystem [14], or transactional checksum [55, 32, 62]. Efforts as transactional write at the filesystem [49, 17, 55, 55, 64] and transactional block device [30, 71, 43, 67, 51] save the application from the overhead of enforcing the storage order associated with filesystem journaling. A school of works address more fundamental aspects in controlling the storage order such as separating the ordering guarantee from durability guarantee [8], providing a programming model to define the ordering dependency among the set of writes [19], persisting a data block only when the result needs to be externally visible [48]. These works share the same essential principle in controlling the storage order, transfer-and-flush. For example, OptFS [8] checkpoints the data blocks only after the associated journal transaction becomes durable. Featherstitch [19] realizes the ordering dependency between the patchgroups via interleaving them with transfer-and-flush.

In this work, we revisit the issue of eliminating the transfer-and-flush overhead in modern IO stack. We aim at developing an IO stack where the host can dispatch the following command before the data blocks associated with the preceding command becomes durable and before the preceding command is serviced and yet the host can enforce the storage order between them.

We develop a Barrier Enabled IO stack which effectively addresses our design objective. Barrier enabled IO stack consists of the cache barrier-aware storage device, the order preserving block device layer and the barrier enabled filesystem. Barrier enabled IO stack is built upon the foundation that the host can control a certain partial order in which the cache contents are flushed, persist order. Different from rotating media, the host can enforce a persist order without the risk of getting anomalous delay in the Flash storage. With reasonable complexity, the storage controller can be made to flush the cache contents satisfying a certain ordering condition from the host [50, 56, 59]. The mobile Flash storage standards already defines “cache barrier” command [28] which precisely serves this purpose. For order preserving block device layer, the command dispatch mechanism and the IO scheduler of the block device layer are overhauled so that they can preserve partial order in the incoming sequence of the requests in scheduling them. For barrier enabled filesystem, we define new interfaces, fbarrier() and fdatabarrier() to exploit the nature of order preserving block device layer. The fbarrier() and the fdatabarrier() system calls are the ordering guarantee only counter part of fsync() and fdatasync(), respectively. fbarrier() shares the same semantics as osync() of OptFS [8]; it writes the dirty pages, triggers filesystem journal commit and returns without persisting them. fdatabarrier() ensures the storage order between its preceding writes and the following writes without flushing the writeback cache in between and without waiting for DMA completion of the preceding writes. It is a storage version of the memory barrier, e.g. mfence [52]. OptFS does not provide the one equivalent to fdatabarrier(). The order-preserving block device layer is filesystem-agnostic. We can implement fbarrier() and fdatabarrier() in any filesystems. We modify EXT4 to support fbarrier() and fdatabarrier(). We only present our result of EXT4 filesystem due to the space limit. We modify the journaling module of EXT4 and develop Dual Mode journaling for order preserving block device. We call the modified version of EXT4, the BarrierFS.

Barrier Enabled IO stack not only removes the flush overhead but also the transfer overhead in enforcing the storage order. While large body of the preceding works successfully eliminate the flush overhead, few works dealt with the overhead of DMA transfer in storage order guarantee. The benefits of Barrier Enabled IO stack include the following:

- The application can control the storage order virtually without any overheads; without being blocked or without stalling the queue.
- The latency of a journal commit decreases significantly. The journaling module can enforce the storage order between the journal logs and the journal commit mark without interleaving them with flush and without interleaving them with DMA transfer.
- Throughput of the filesystem journaling improves significantly. Dual Mode journaling commits multi-

1 The source codes are currently unavailable to public to abide by the double blind rule of the submission. We plan to open-source it shortly.
ple transactions concurrently and yet can guarantee the durability of the individual journal commit.

Eliminating all the inefficiencies, the host now can successfully exploit the concurrency and the parallelism in the underlying storage satisfying all ordering constraints. Relaxing the durability of a transaction, SQLite performance and MySQL performance increase as much as by 73× and by 43×, respectively, in server storage.

The rest of the paper is organized as follows. Section 2 introduces the background. Section 3, section 4 and section 5 explain the block device layer, the filesystem layer, and the application of Barrier Enabled IO stack, respectively. Section 6 and section 7 discuss the result of the experiment and surveys the related works, respectively. Section 8 concludes the paper.

2 Background

2.1 Orders in IO stack

A write request travels a complicated route until the associated data blocks reach the storage surface. The filesystem puts the request to the IO scheduler queue. The block device driver removes one or more requests from the queue and constructs a command. It probes the device and dispatches the command if the device is available. The device is available if the command queue at the storage device is not full. Arriving at the storage device, the command is inserted into the command queue. The storage controller removes the command from the command queue and services it, i.e. transfers the data block between the host and the storage. When the transfer finishes, the device sends the completion signal to the host. The contents of the writeback cache are committed to storage surface either periodically or by an explicit request from the host.

We define four types of orders in the IO stack; Issue Order, \( I \), Dispatch Order, \( D \), Transfer Order, \( C \), and Persist Order, \( P \). The issue order \( I = \{i_1, i_2, \ldots, i_n\} \) is a set of write requests issued by the application or by the file system. The subscript denotes the order in which the requests enter the IO scheduler. The dispatch order \( D = \{d_1, d_2, \ldots, d_n\} \) denotes a set of the write requests which are dispatched to the storage device. The subscript denotes the order in which the requests leaves the IO scheduler. Transfer order, \( C = \{c_1, c_2, \ldots, c_n\} \), is the set of transfer completions. Persist Order \( P = \{p_1, p_2, \ldots, p_n\} \) is a set of operations which make the associated data blocks durable. Fig. 2 schematically illustrates the layers and the associated orders in the IO stack. We say a certain partial order is preserved if the relative position of the requests against a certain designated request, barrier, are preserved. We use the notation ‘=’ to denote that a certain partial order is preserved. We briefly summarize the source of arbitration at each layer.

- \( I \neq D \). IO scheduler reorders and coalesces the IO requests subject to their optimization criteria, e.g. CFQ, DEADLINE, etc. When there is no scheduling mechanism, e.g. NO-OP scheduler or NVMe interface, the dispatch order may be equal to the issue order.
- \( D \neq C \). Storage controller freely schedules the commands in its command queue. Also, the data blocks can be transferred out of order due to the errors, time-out and retry.
- \( C \neq P \). The cache replacement algorithm, mapping table update algorithm, and storage controller’s policy to schedule Flash operations governs the persist order independent of the order in which the data blocks are transferred.

Due to all these sources of arbitrations, the modern IO stack is said to be orderless [7].

2.2 Transfer-and-Flush

Enforcing a storage order corresponds to preserving a partial order between issue order \( I \) and persist order \( P \), i.e. satisfying the condition \( I = P \). It is equivalent to collectively enforcing the individual ordering constraints between the layers:

\[
(I = P) \equiv (I = D) \land (D = C) \land (C = P) \quad (1)
\]

Modern IO stack has evolved under the assumption that the host cannot control the persist order, i.e. \( C \neq P \). Persist order specifically denotes the order in which the contents in the writeback cache are persisted whereas storage order denotes an order in which the write requests from the filesystem are persisted. For rotating media such as hard disk drive, the disk scheduling is entirely left to the storage device due to its complicated sector geometry hidden from outside [20]. Blindly enforcing a certain persist order may bring unexpected delay in IO service. Inability to control the persist order, \( C \neq P \), is a fundamental limitation of the modern IO stack, which makes the condition \( I = P \) in Eq. 1 unsatisfiable.
To circumvent this limitation in satisfying a storage order, the host takes the indirect and expensive resort to satisfy each component in Eq. 1. First, after dispatching the write command to the storage device, the caller is blocked until the associated DMA transfer completes, Wait-on-Transfer. This is to prohibit the storage controller from servicing the commands in out-of-order manner and to satisfy the transfer order, $D = C$. This may stall the command queue. When the DMA transfer completes, the caller issues the flush command and blocks again waiting for its completion. When the flush returns, the caller wakes up and issues the following command; Wait-on-Flush. These two are used in tandem leaving the caller under a number of context switches. Transfer-and-flush is unfortunate sole resort in enforcing the storage order in a modern orderless IO stack.

### 2.3 Analysis: fsync() in EXT4

We examine how the EXT4 filesystem controls the storage order among the data blocks, journal descriptor, journal logs and journal commit block in fsync() in Ordered mode journaling. In Ordered mode, EXT4 ensures that data blocks are persisted before the associated journal transaction does.

Fig. 3 illustrates the behavior of an fsync(). The application dispatches the write requests for the dirty pages, $D$. After dispatching the write requests, the application blocks and waits for the completion of the associated DMA transfer. When the DMA transfer completes, the application thread resumes and triggers the JBD thread to commit the journal transaction. After triggering the JBD thread, the application thread sleeps again. When the JBD thread makes journal transaction durable, the fsync() returns, waking up the caller. The JBD thread should be triggered only after $D$ are completely. Otherwise, the storage controller may service the write requests for $D$, $JD$ and $JC$ in out-of-order manner and storage controller may persist the journal transaction prematurely before $D$ reaches the writeback cache. In this happens, the filesystem can be recovered incorrectly in case of the unexpected system failure.

A journal transaction consists of the journal descriptor block, one or more log blocks and the journal commit block. A transaction is usually written to the storage with two requests: one for writing the coalesced chunk of the journal descriptor block and the log blocks and the other for writing the commit block. In the rest of the paper, we will use $JD$ and $JC$ to denote the coalesced chunk of the journal descriptor and the log blocks, and the commit block, respectively. JBD needs to enforce the storage order in two situations. $JD$ needs to be made durable before $JC$. The journal transactions need to be made durable in the order in which they have been committed. When any

$$D \rightarrow xfer \rightarrow JD \rightarrow xfer \rightarrow JC \rightarrow xfer \rightarrow flush$$ (2)  

In early days, the block device layer was responsible for issuing the flush and for waiting for its completion [63][8]. This approach blocks not only the caller but all the other requests which share the same dispatch queue [14]. Since Linux 2.6.37 kernel, this role has been migrated from the block device layer to the filesystem layer [15]. The filesystem uses flush option (REQ_FLUSH) and force-unit-atomic option (REQ_FUA) in writing $JC$ and the filesystem blocks until it completes. With FLUSH option, the storage device flushes the writeback cache before servicing the command. With FUA option, the storage controller writes a given block directly to the storage surface. The last four steps in Eq. 2 can be compressed into a write request with FLUSH|FUA option. When the filesystem is responsible for waiting for the completion of Flash, the other commands in the dispatch queue can progress after $JC$ is dispatched. In both approaches, the caller is subject to transfer-and-flush overhead to interleave $JD$ and $JC$.

### 3 Order Preserving Block Device Layer

#### 3.1 Design

We overhaul the IO scheduler, the dispatch module and the write command to satisfy each of three conditions, $I = D$, $D = C$, and $C = P$, respectively.
In the legacy IO stack, the host has been entirely responsible for controlling the storage order; the host postpones sending the following command until it ensures that the result of the preceding command is made durable. In Barrier enabled IO stack, the host and the storage device share the responsibility. The host side block device layer is responsible for dispatching the commands in order. The host and the storage device collaborate with each other to transfer the data blocks (or to service the commands, equivalently) in order. The way in which the host and the storage device collaborate with each other will be detailed shortly. The storage device is responsible for making them durable in order. This effective orchestration between the host and the storage device saves the IO stack from the overhead of transfer-and-flush based storage order guarantee. Fig. 4 illustrates the organization of Barrier Enabled IO stack.

The order preserving block device layer is responsible for dispatching the commands in order and for having them serviced in order. The IO scheduler and the command dispatch module is redesigned to preserve the order. Order preserving block device layer defines two types of write requests: orderless and order-preserving. There exists special type of order-preserving request called barrier. We introduce two new attributes REQ.ORDERED and REQ.BARRIER for the order-preserving request and the barrier request, respectively. We call a set of order-preserving write requests which can be reordered with each other as an epoch [13]. A barrier request is used to delimit an epoch.

3.2 barrier write, the command

The “cache barrier”, or “barrier” for short, command is defined in the standard command set for mobile Flash storage [28]. When the storage controller receives the barrier command, the controller guarantees that the data blocks transferred following the barrier command reach the storage surface after the data blocks transferred before the barrier command do without flushing the cache in between. A few eMMC products in the market support cache barrier command [11, 12]. Via barrier command, the

IO stack can satisfy the persist order without cache flush. The essential condition $\mathcal{E} = \emptyset$ in ensuring the storage order can now be satisfied with the barrier command.

We start our effort with devising a more efficient barrier write command. Implementing a barrier as a separate command occupies one entry in the command queue and costs the host the latency of dispatching a command. To avoid this overhead, we define a barrier as a command flag, REQ.BARRIER, to the write command as in the case of REQ.FUA or REQ.FLUUSH. In our implementation, we designate one unused bit in the SCSI command as a barrier flag.

We discuss the implementation aspect of a barrier command. It is a matter of how the storage controller can enforce the persist order imposed by the barrier command. When the Flash storage device has Power Loss Protection (PLP) feature, e.g. supercapacitor, supporting a barrier command is trivial. Thanks to PLP, the writeback cache contents are always guaranteed to be durable. The storage controller can flush the writeback cache in any order fully utilizing its parallelism and yet can guarantee the persist order. There is no performance overhead in enforcing the persist order.

For the devices without PLP, the barrier command can be supported in three ways: in-order write-back, transactional write-back or in-order recovery from crash. In in-order write-back, the storage controller flushes data blocks in epoch basis and inserts some delay in between if necessary. It may fail to fully exploit the underlying parallelism in the storage controller. In transactional write, the storage controller flushes the writeback cache contents as a single atomic unit [56, 39]. Since all epochs in the writeback cache are are flushed together, the constraint imposed by the barrier command is well satisfied. The performance overhead of transactional flush is 12% in worst case with a traditional commit approach but can be eliminated by maintaining next page pointer at the spare area of the Flash page [56].

The in-order recovery method guarantees the persist order imposed by the barrier command through crash recovery routine. When multiple controller cores concurrently write the data blocks to multiple channels, one may have to use sophisticated crash recovery protocol such as ARIES protocol [45] to recover the storage to consistent state. If the entire Flash storage is treated as a single log device, we can use simple crash recovery algorithm used in LFS [59]. Since the persist order is enforced by the crash recovery logic, the controller is able to flush the writeback cache as if there is no ordering dependency. The controller is saved from performance penalty at the cost of complexity in the recovery routine.

We implement the cache barrier command in UFS device, which is a commercial product used in the smartphone. We use simple LFS style recovery routine. The
3.3 Epoch Based IO scheduling

There are three scheduling principles in Epoch based IO scheduling. First, it preserves the partial order between the epochs. Second, the requests within an epoch can be freely scheduled with each other. Third, the orderless requests can be scheduled freely across the epochs. It satisfies $S = D$ condition.

The Epoch Based IO scheduler uses existing IO scheduler, e.g. CFQ, NO-OP and etc., to schedule the IO requests within an epoch. The key ingredient of the Order Preserving IO scheduler is Epoch Based Barrier Reassignment. When the IO request enters the scheduler queue, the order preserving IO scheduler examines if it is a barrier request. If the request is not a barrier request, it is inserted as normal requests. If the request is a barrier write request, IO scheduler removes the barrier flag from the request and inserts it to the queue. After the scheduler inserts a barrier write, the scheduler stops accepting more requests. The IO scheduler re-orders and merges the IO requests in the queue based upon its own scheduling discipline e.g. FIFO, SCAN, CFQ. The requests in the queue either are orderless or belong to the same epoch. Therefore, they can be freely scheduled with each other without violating the ordering condition. The merged request will be order-preserving if one of the constituents is order-preserving. The IO scheduler designates the order-preserving request that leaves the queue last as a new barrier. This mechanism is called Epoch Based Barrier Reassignment. When there is no more order-preserving requests in the queue, the IO scheduler starts accepting the IO requests. If the IO scheduler unblocks the queue, there can be one or more orderless requests in the queue. These orderless requests can be scheduled with the other requests in the following epoch. Differentiating the order-preserving requests from orderless ones, we avoid imposing unnecessary ordering constraint on the requests. Currently, the Epoch based IO scheduler is implemented on top of existing CFQ scheduler. Each process defines its own scheduler queue.

Fig. 5 illustrates how the barrier reassignment works. The circular and the rectangular write request denote the order-preserving attribute and barrier attribute, respectively. In Fig. 5, the application calls $fsync()$ and in the mean time, $pdflush$ daemon flushes the dirty pages. In Fig. 5 $fsync()$ creates three write requests: $w_1, w_2$ and $w_4$. The filesystem marks the three requests as ordering preserving ones. The filesystem designates the last request, $w_4$, as a barrier write. $pdflush$ creates three write requests $w_3, w_5$ and $w_6$. They are all orderless. The requests from the two threads are fed to the IO scheduler with $w_1, w_2, w_3, w_5, w_4^{barrier}, w_6$ in order. When the barrier write, $w_4$, enters the queue, the scheduler stops accepting the new request. There are only five requests in the queue, $w_1, w_2, w_3, w_4$ and $w_5$. $w_6$ cannot be inserted at the queue since the queue is blocked. The IO scheduler reorders the them and dispatches them in $w_2, w_3, w_4, w_5, w_1$ order. After they are scheduled, $w_1$ leaves the queue last. The IO scheduler puts the barrier flag to $w_1$. In this scenario, the request $w_6$ is going to be scheduled with the requests in the following epoch.

3.4 Order Preserving Dispatch

The order preserving dispatch is a fundamental innovation of this work. In order preserving dispatch, the host dispatches the following write request when the storage device acknowledges that the preceding request has successfully been received (6(a)) and yet the transfer order between the two requests are preserved, i.e. $S = C$. The order preserving dispatch guarantees the transfer order without blocking the caller. Legacy IO stack controls the transfer order with Wait-On-Transfer. Wait-On-Transfer not only exposes the caller to the context switch overhead but also makes the IO latency less predictable. It may stall the storage device since the caller postpones dispatching the following command till the preceding command is serviced. Order preserving dispatch eliminates
all these overheads.

For order preserving dispatch, the only thing the host block device driver does is to set the priority of a barrier write command to *ordered* when dispatching it. Then, the SCSI compliant storage device automatically guarantees the transfer order constraint in serving the requests. SCSI standard defines three command priority levels: *head of the queue*, *ordered*, and *simple* [57], with which the incoming command is put at the head of the command queue, tail of the command queue or at arbitrary position determined by the storage controller. In addition, the simple command cannot be inserted in front of the existing "ordered" or "head of the queue" commands. The *head of the queue* priority is used when a command requires an immediate service, e.g. flush command. Via setting the priority of barrier write command to *ordered*, the host ensures the the data blocks associated with the write requests in the preceding epoch are transferred ahead of the data blocks associated with the barrier write. Likewise, the data blocks associated with the following epoch are transferred after the data blocks associated with the barrier write is transferred. The transfer order condition is satisfied.

The caller may be blocked after dispatching the write request. This can happen when the device is unavailable or the caller is switched out involuntarily, e.g. time quantum expires. For both cases, the block device driver of the order preserving dispatch module uses the same error handling routine adopted by the existing block device driver; the kernel daemon inherits the task and retries dispatching the request after a certain time interval, e.g., 3 msec for SCSI device [57] (Fig. 6(b)). The thread resumes once the request is dispatched successfully.

4 BarrierFS: Barrier Enabled Filesystem

4.1 Programming Model

We propose two new filesystem interfaces, `fbarrier()` and `fdatabarrier()` which are the ordering guarantee only counter part to `fsync()` and `fdatasync()`, respectively. `fbarrier()` shares the same semantics with `osync()` in OptFS [8]. The salient feature of BarrierFS is `fdatabarrier()`.After dispatching the write requests for dirty pages. With `fdatabarrier()`, the application can enforce a storage order virtually without any overhead; without flush, without waiting for DMA completion and even without context switch. The following codelet illustrates the usage of the `fdatabarrier()`.

```c
write(fileA, "Hello")
fdatabarrier(fileA);
write(fileA, "World")
```

It ensures that “Hello” is written to the storage surface ahead of “World”. Modern applications have been using expensive `fdatasync()` to guarantee both durability and ordering. For example, SQLite which is the default DBMS in mobile device, such as Android, iOS or Tizen uses `fdatasync()` to ensure that the updated database node reach the disk surface ahead of the updated database header. In SQLite, `fdatabarrier()` can replace the `fdatasync()` when it is used for ensuring the storage order, not the durability.

The Barrier Enabled IO stack is filesystem agnostic. `fbarrier()` and `fdatabarrier()` can be implemented in any filesystem using proposed order preserving block device layer. As a seminal work, we modify the EXT4 filesystem for order preserving block device layer. We optimize `fsync()` and `fdatasync()` for order preserving block device layer and newly implement `fbarrier()` and `fdatabarrier()`. We name the modified EXT4 as BarrierFS. `fbarrier()` in BarrierFS supports all journal modes in EXT4; WRITEBACK, ORDERED and DATA.

4.2 Dual Mode Journaling

Committing a journal transaction essentially consists of two separate tasks: dispatching write commands for JD and JC to the storage (host side) and making them durable (storage side). In the order preserving block device design, the host (the block device layer) is responsible for controlling the dispatch order and transfer order while the storage controller takes care of handling the persist order. The design of order preserving block device layer naturally supports separation of the control plane (dispatching the write requests) and the data plane
The flush thread flushes the cache, removes the associating thread tags both requests with \texttt{REQ} and \texttt{FUA} respectively. This mechanism is called Dual Mode Journaling. For effective separation, these two planes should work independently with minimum dependency. For filesystem journaling, we allocate separate threads for dispatching the write requests and for making them durable: commit thread and flush thread, respectively. This mechanism is called Dual Mode Journaling.

The commit thread is responsible for dispatching the write requests for \( JD \) and \( JC \). In BarrierFS, the commit thread tags both requests with \texttt{REQ} and \texttt{REQ\_BARRIER} so that \( JD \) and \( JC \) are transferred and are guaranteed to be persisted in order. After the dispatching write request for \( JC \), the commit thread inserts the journal transaction to the committing transaction list. In ordering guarantee (\texttt{fbarrier()}), the commit thread wakes up the caller. In the legacy IO stack, JBD thread interleaves the write request for \( JC \) and \( JD \) with transfer-and-flush. In BarrierFS, the commit thread dispatches them in order-preserving dispatch discipline without Wait-For-Transfer overhead and with Wait-For-Flush overhead.

The flush thread is responsible for (i) issuing the flush command, (ii) handling error and retry and (iii) removing the transaction from the committing transaction list. The flush thread is triggered when the \( JC \) is transferred. If the journaling is triggered by \texttt{fbarrier()}, the flush thread removes the transaction from the committing transaction list and returns. It does not call flush. There is no caller to wake up. If the journaling is initiated by \texttt{fsync()}, the flush thread flushes the cache, removes the associated transaction from the committing transaction list and wakes up the caller. Via separating the control plane (commit thread) and data plane (flush thread), the commit thread can commit the following transaction after it is done with dispatching the write requests for preceding journal commit. In Dual Mode journaling, there can be more than one committing transactions in flight.

In \texttt{fsync()} or \texttt{fbarrier()}, the BarrierFS dispatches the write request for \( D \) as an order-preserving request. Then, the commit thread dispatches the write request for \( JD \) and \( JC \) both with order-preserving and barrier write. As a result, \( D \) and \( JD \) form a single epoch while \( JC \) by itself forms another. A journal commit consists of the two epoches: \{\( D, JD \)\} and \{\( JC \)\}. An \texttt{fsync()} in barrierFS can be represented as in Eq. 3. Eq. 3 also denotes the \texttt{fbarrier()}.

\begin{equation}
D \rightarrow JD_{BAR} \rightarrow JC_{BAR} \rightarrow xfer \rightarrow flush \quad (3)
\end{equation}

The benefit of Dual Mode Journaling is substantial. In EXT4 (Fig. 7(a)), an \texttt{fsync()} consists of a tandem of three DMA's and two flushes interleaved with context switches. In BarrierFS, an \texttt{fsync()} consists of single flush, three DMA's(Fig. 7(b)) and fewer number of context switches. The transfer-and-flush between \( JD \) and \( JC \) are completely eliminated. \texttt{fbarrier()} returns almost instantly after the commit thread dispatches the write request for \( JC \).

BarrierFS forces journal commit if \texttt{fdatasync()} or \texttt{fdatabarrier()} do not find any dirty pages. Through this scheme, \texttt{fdatasync()} (or \texttt{fdatabarrier()}) can delimit an epoch despite the absence of the dirty pages.

\subsection{Multi-Transaction Page Conflict}

A buffer page can belong to only one journal transaction at a time. Blindly inserting a buffer page to the running transaction may yield removing it from the committing transaction before it becomes durable. We call this situation as page conflict. In both EXT4 and BarrierFS, when the application thread inserts a buffer page to the running transaction, it checks if the buffer page is being held by the committing transaction. If so, the application blocks without inserting it to the running transaction. When the JBD thread of EXT4 (or flush thread in BarrierFS) has made the committing transaction durable, it identifies the conflict pages in the committed transaction and inserts them to the running transaction. In EXT4, there is only one committing transaction at a time. The running transaction is guaranteed to be conflict free when the JBD thread resolves the page conflicts from the committed transaction. In BarrierFS, the running transaction can conflict with more than one
committing transactions, multi-transaction page conflict. When the flush thread resolves the page conflicts from a committed transaction, the running transaction may still conflict with the other committing transactions. If the running transaction is committed prematurely with conflicted pages missing, the storage order can be compromised. Whenever the flush thread resolves the page conflicts and notifies the commit thread about its completion of persisting a transaction, the commit thread has to scan all the pages in the other committing transactions for page conflict. To reduce the overhead of scanning the pages, we introduce conflict-page list. The application thread inserts the buffer page to the conflict-page list if the buffer page is being held by one of the committing transactions. When the flush thread has made the committing transaction durable, the flush thread inserts the conflict pages to the buffer page list of the running transaction and removes them from the conflict-page list. The commit thread can start committing a running transaction only when conflict-page list is empty.

4.4 Analysis

We examine how the journaling throughput may vary subject to different methods of journal commit: BarrierFS, EXT4 with no-barrier option, EXT4 with supercap SSD and and plain EXT4. Fig. 8 schematically illustrates the behaviors. With no-barrier mount option, filesystem does not issue flush command in fsync() or fdatasync(). $t_D$, $t_C$, and $t_F$ denote the dispatch latency, transfer latency, and flush latency associated with committing a journal transaction, respectively. In particular, $t_F$ denotes the total flush latency in supercap SSD.

With supercap SSD, EXT4 (quick flush), the journal commits are interleaved by $t_D + t_C + t_F$. The host observes the round-trip delay of the flush command and the associated context switch overhead, $t_C$. $t_F$ is not negligible in Flash storage. EXT4 with no-barrier option, EXT4 (no flush), can commit a new transaction once all the associated blocks are transferred to the storage. The journaling is interleaved by command dispatch and DMA transfer, $t_D + t_C$. In BarrierFS, the commit thread keeps dispatching the journal commit operations without waiting for the completion of the transfer. The interval between the successive journal commit can be as small as $t_D$.

Fsync() accounts for dominant fraction of IO in modern applications, e.g. mail server [50] or OLTP. 90% of IO’s in the TPC-C workload is created by fsync() for synchronizing the logs to the storage [50]. The order preserving IO stack can significantly improve the performance in these workloads. SQLite can be the application which the Barrier Enabled IO stack benefits the most. SQLite uses fdatasync() not only to guarantee the durability of a transaction but also to control the storage order in various occasions, e.g. between writing the undo-log and storing the journal header and between writing updated database node and writing the commit block $t_C$. In a single insert transaction, SQLite calls fdatasync() four times, three of which are to control the storage order. We can replace them with fsync()’s without compromising the durability of a transaction. Some applications prefer to trade the durability and freshness of the result with the performance and scalability of the operation $t_F$. The benefit of BarrierFS can be more than significant in these applications. One can replace all fsync() and fdatasync() with ordering guarantee counterparts, fdatabarrier() and fdatabarrier(), respectively.

6 Experiment

6.1 Setup

We implement Barrier Enabled IO stack on three different platforms: smartphone (Galaxy S6, Android 5.0.2, Linux 3.10), PC server (4 cores, Linux 3.10.61) and enterprise server (16 cores, Linux 3.10.61). We test three storage devices: mobile storage (UFS 2.0, QD=16, single channel), 850 PRO for server (SATA 3.0, QD=32, 8 channels), 843TN for server (SATA 3.0, QD=32, 8 channels, supercap). We call each of these as UFS, plain-SSD and supercap-SSD, respectively. We implement barrier write command in UFS device. In plain-SSD, we introduce 5% performance penalty to simulate the barrier overhead. For supercap-SSD, we assume that there is no barrier overhead.

2 QD: queue depth
6.2 Order Preserving Block Layer

We examine the performance of 4 KByte random write with different ways of enforcing the storage order. Fig. 9 illustrates the result. In scenario ‘X’ where ‘X’ denotes Wait-On-Transfer, the host sends the following request after the data block associated with the preceding request is completely transferred. Despite the absence of the flush overhead, the storage devices exhibit less than 50% of its plain buffered write performance, the scenario ‘P’. All three devices are severely underutilized. Average queue depths in all three devices are less than one. Wait-On-Transfer overhead in modern IO stack prohibits the host from properly exploiting the underlying Flash storage. In scenario ‘B’ where ‘B’ denotes Barrier, the IO performance increases at least by 2× against scenario ‘X’. The average queue depths reach near the maximum in all three Flash storages. An fdatabarrier() is not entirely free. We observe 1% to 25% performance deficiency when it is compared against the plain buffered write. Plain buffered write exhibits shorter queue depth than barrier write does (Fig. 9). This is because in plain buffered write, the IO scheduler merges the multiple requests and the number of commands dispatched to the storage device decreases.

Fig. 10 is another manifestation of fdatabarrier(). The storage performance is closely related to the command queue utilization [5]. When the requests are interleaved with DMA transfer, the queue depth never goes beyond one (Fig. 10(a) and Fig. 10(c)). When the write request is followed by fdatabarrier(), the queue depth grows near to its maximum in all three storage. (Fig. 10(b) and Fig. 10(d)). Order preserving block layer enables the host to fully exploit the concurrency and the parallelism of the underlying Flash storage.

6.3 Filesystem Journaling

Latency: In plain-SSD and supercap-SSD, the average ffsync() latency decreases by 40% when we use BarrierFS against when we use EXT4 (Table 1). UFS experiences more significant reduction in ffsync() latency than the SSD’s do. The smartphone uses transactional checksum in filesystem journaling. With BarrierFS, we can eliminate not only the transfer overhead but also the checksum overhead. The ffsync() latency decreases by 60% in BarrierFS. In supercap-SSD and UFS, the ffsync() latencies at 99.99th percentile are 30× of the average ffsync() latency (Table 1). Using BarrierFS, the tail latencies at 99.99th percentile decrease by 50%, 20% and 70% in UFS, plain-SSD and supercap-SSD, respectively, against EXT4.

| (%) | UFS | plain-SSD | supercap-SSD |
|-----|-----|-----------|--------------|
| μ   | 1.29 | 0.31      | 0.51         |
| Median | 1.20 | 0.44      | 0.40         |
| 99.9% | 4.15 | 3.51      | 11.41        |
| 99.99% | 22.83 | 9.02      | 16.09        |
| 99.99% | 33.10 | 17.60     | 17.26        |

Table 1: ffsync() latency statistics (msec)

Context Switches: We examine the number of application level context switches in various modes of
Figure 12: Queue Depth Changes in BarrierFS: write() followed by fsync() vs. write() followed by fbarrier() journaling. Fig. 11 illustrates the result. In EXT4-DR, fsync() wakes up the caller twice; after DMA transfer of D completes and after the journal transaction is made durable. This applies to all three Flash storages. In BarrierFS, fsync() wakes up the caller only once; after the transaction is made durable. In UFS and supercap SSD, fsync() of BFS-DR wakes up the caller twice in entirely different reasons. In UFS and supercap-SSD, the interval between the successive write requests are much smaller than the timer interrupt interval due to small flush latency. As a result, write() requests rarely update the time fields of the inode and fsync() becomes an fdatasync(). fdatasync() wakes up the caller twice in BarrierFS; after transferring D and after flush completes. The plain-SSD uses TLC flash. The interval between the successive write()’s can be longer than the timer interrupt interval. In plain-SSD, fsync() occasionally commits journal transaction and the average number of context switches becomes less than two in BFS-DR for plain-SSD.

BFS-OD manifests the benefits of BarrierFS. The fbarrier() rarely finds updated metadata since it returns quickly. Most fbarrier() calls are serviced as fdatabarrier(). fdatabarrier() does not block the caller and it does not release CPU voluntarily. The number of context switches in fbarrier() is much smaller than EXT4-OD. BarrierFS significant improves the context switch overhead against EXT4.

Command Queue Utilization: In BarrierFS, fsync() drives the queue up to three (Fig. 12(a)). Theoretically, it can drive the queue depth up to three because the host can dispatch the write requests for D, JD and JC, in tandem. According to our instrumentation, there exists 160 µsec context switch interval between the application thread and the commit thread. It takes approximately 70µsec to transfer a 4 KByte block from the host to device cache. The command from the application thread is serviced before the commit thread dispatches the command for writing JD. In fbarrier(), BarrierFS successfully saturates the command queue (Fig. 12(b)). The queue depth increases to fifteen.

Throughput: We examine the throughput of filesystem journaling under varying number of CPU cores. We use modified DWSL workload in fxmark [44]. In DWSL workload, each thread performs 4 Kbyte allocating write followed by fsync(). Each thread operates on its own file. Each thread writes total 1 GByte. BarrierFS exhibits much more scalable behavior than EXT4 (Fig. 13). In plain-SSD, BarrierFS exhibits 2× performance against EXT4 in all numbers of cores (Fig. 13(a)). In supercap-SSD, the performance saturates with six cores in both EXT4 and BarrierFS. BarrierFS exhibits 1.3× journaling throughput against EXT4 at the full throttle (Fig. 13(b)).

6.4 Mobile Workload: SQLite

In mobile storage, BarrierFS achieves 75% performance improvement against EXT4 in default PERSIST journal mode under durability guarantee (Fig. 14). We replace first three fdatasync()’s with fdatabarrier()’s among all four fdatasync()’s in a transaction. We keep the last fdatasync() for the durability of a transaction. In Ordering guarantee, we replace all four fdatasync()’s with fdatabarrier()’s. When we remove the durability requirement, the performance increases by 2.8× in PERSIST mode against the baseline EXT4. In WAL mode, SQLite issues fdatasync() once in every commit and there is not much room for improvement for BarrierFS.

The benefit of eliminating the Transfer-and-flush is more significant as the storage has higher degree of parallelism and slow Flash device. In plain-SSD, SQLite exhibits 73× performance gain in BFS-OD against baseline EXT4-DR.

6.5 Server Workload

We run two workloads: varmail workload in FILEBENCH [68] and OLTP-insert workloads from sysbench [44]. Sysbench is database workload and uses MySQL [46]. varmail is metadata intensive workload. We also test OptFS [3]. We use osync() in OptFS.

We perform two sets of experiments. First, we leave the application intact and replace the EXT4 with BarrierFS (EXT4-DR and BFS-DR). We compare the
fsync() performance between BarrierFS and EXT4. The second set of experiment is for ordering guarantee. In EXT4, we use nobarrier mount option. In BarrierFS, we replace fsync() with fbarrier(). Fig. 15 illustrates the result.

In plain-SSD, BFS-DR brings 60% performance gain against EXT4-DR in varmail workload. This is due to the more efficient implementation of fsync() in BarrierFS. The benefit of BarrierFS manifests itself when we relax the durability guarantee. The varmail workload is known for its heavy fsync() traffic. In EXT4-OD, the journal commit operations are interleaved by DMA transfer latency. In BFS-OD, the journal commit operations are interleaved by the dispatch latency. The Dual mode journal can significantly improve the journaling throughput via increasing the concurrency in journal commit. With ordering guarantee, BarrierFS achieves 80% performance gain against EXT4 with no-barrier option.

In MySQL, BFS-OD prevails EXT4-OD, by 12%. The performance increases 43× when we replace the fsync() of EXT4 with fbarrier().

Notes on OptFS: In SQLite (Fig. 14(b)), varmail and MySQL (Fig. 15), we observe that OptFS does not show as good performance in Flash storage as it does in the rotating media [8]. OptFS is elaborately designed to reduce the seek overhead inherent in Ordered mode journaling of EXT4. OptFS achieves this objective via two innovations: via flushing larger number of transactions together and via selectively journaling the data blocks. Benefit of eliminating a seek overhead is marginal for Flash storage. Due to this reason, in varmail workload which rarely entails selective data mode journaling, OptFS and EXT4-OD exhibit similar performance in Flash storage(Fig. 15). The selective data mode journaling increases the amount of pages to scan for osync(), only a few of which can be dispatched to the storage. The selective data mode journaling can negatively interfere with the osync() especially when the underlying storage has short latency. In [8], MySQL performance decreases to one thirds in OptFS against EXT4-OD and the selective data mode journaling has been designated as its prime cause. Our MySQL workload creates even larger amount of selective data journaling and the performance of OptFS corresponds to one eights of that of EXT-OD under MySQL workload (Fig. 15).

7 Related Work

OptFS [8] is the closest work of our sort; they proposed a new journaling primitive osync() which returns without persisting the journaling transaction and yet which guarantees that the write requests associated with journal commits are stored in order. OptFS does not provide the filesystem primitive that corresponds to fdbarrier() in our Barrier Enabled IO stack. osync() still relies on Wait-On-Transfer in enforcing the storage order. Featherstitch[19] propose a programming model to specify the set of requests that can be scheduled together. patchgroup and the ordering dependency between them pg_depend(). While xsyncfs [48] successfully mitigates the overhead of fsync(), xsyncfs maintains complex causal dependencies among buffered updates. An order preserving block device layer can make the implementation of xsyncfs much simpler. NoFS (no order file system) [9] introduces “backpointer” to entirely eliminate the transfer-and-flush ordering requirement in the file system. However, it does not support atomic transactions.

A few works proposed to use multiple running transaction or multiple committing transaction to circumvent the transfer-and-flush overhead in filesystem journaling [48, 29, 54], to improve journaling performance or to isolate errors. IceFS [48] allocates separate running transactions for each container. SpanFS [29] splits a journal region into multiple partitions and allocates committing transactions for each partition. CCFS [54] allocates separate running transactions for individual threads. These systems, where each journaling session still relies on the transfer-and-flush mechanism in enforcing the intra- and inter-transaction storage orders, are complementary to our work.

A number of file systems provide a multi-block atomic write feature [17, 33, 53, 66] to relieve applications from
the overhead of logging and journaling. These file systems internally use the transfer-and-flush mechanism to enforce the storage order between write requests for data blocks and associated metadata. An order preserving block device can effectively mitigate overheads incurred when enforcing the storage order in these file systems.

8 Conclusion

In this work, we develop an Barrier Enabled IO stack to address the transfer-and-flush overhead inherent in the legacy IO stack. Barrier Enabled IO stack effectively eliminates the transfer-and-flush overhead associated with controlling the storage order and is successful in fully exploiting the underlying Flash storage. We like to conclude this paper with two important observations. First, “cache barrier” is a necessity and a luxury. “cache barrier” is an essential tool for the host to control the persist order which has not been possible before. Currently, cache barrier command is only available in the standard command set for mobile storage. Given its implication on IO stack, it should be available in all range of the storage device ranging from the mobile storage to the high performance Flash storage with supercap. Second, eliminating a “Wait-On-Transfer” overhead is not an option. It blocks the caller and stalls the command queue leaving the storage device being severely underutilized. As the storage latency becomes shorter, the relative cost of “Wait-On-Transfer” can become more significant.

Despite all the preceding sophisticated techniques to optimize the legacy IO stack for Flash storage, we carefully argue that the IO stack is still fundamentally driven by the old legacy that the host cannot control the persist order. This work shows how the IO stack can evolve when the persist order can be controlled and its substantial benefit. We hope that this work serves as a possible basis for the future IO stack in the era of Flash storage.

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conference proceedings, journal articles, and technical reports related to computer science, specifically focusing on topics like flash storage, file systems, and data structures.

For example, a common theme is the optimization of file systems for smartphones, as seen in the reference to "Optimization for Smartphones. In Proc. of USENIX ATC 2013 (San Jose, CA, USA, Jun 2013)." This indicates a discussion on improving file system performance for mobile devices.

Another example is the use of flash memory in modern hardware, as highlighted in "A 128Gb 2b/cell NAND flash memory in 14nm technology with the journaling of journal anomaly. In Proc. of IEEE ICCD 2013." This suggests a focus on the cutting-edge technology and its implications for file system design.

The references also cover topics such as transactional storage, as seen in "A lightweight transactional design in flash-based ssds to support crash consistency. In Proc. of USENIX ATC 2016." This points to advancements in maintaining data integrity in volatile environments.

Overall, the text reflects the rapid evolution of computer science, particularly in the areas of storage and file systems, with a strong emphasis on practical applications and theoretical underpinnings.
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