ZZ and the Art of Practical BFT Execution

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Abstract

The high replication cost of Byzantine fault-tolerance (BFT) methods has been a major barrier to their widespread adoption in commercial distributed applications. We present ZZ, a new approach that reduces the replication cost of BFT services from $2f + 1$ to practically $f + 1$. The key insight in ZZ is to use $f + 1$ execution replicas in the normal case and to activate additional replicas only upon failures. In data centers where multiple applications share a physical server, ZZ reduces the aggregate number of execution replicas running in the data center, improving throughput and response times. ZZ relies on virtualization—a technology already employed in modern data centers—for fast replica activation upon failures, and enables newly activated replicas to immediately begin processing requests by fetching state on-demand. A prototype implementation of ZZ using the BASE library and Xen shows that, when compared to a system with $2f + 1$ replicas, our approach yields lower response times and up to 33% higher throughput in a prototype data center with four BFT web applications. We also show that ZZ can handle simultaneous failures and achieve sub-second recovery.

1 Introduction

Today's enterprises rely on data centers to run their critical business applications. As users have become increasingly dependent on online services, malfunctions have become highly problematic, resulting in financial losses, negative publicity, or frustrated users. Consequently, maintaining high availability of critical services is a pressing need as well as a challenge in modern data centers.

Byzantine fault tolerance (BFT) is a powerful replication approach for constructing highly-available services that can tolerate arbitrary (Byzantine) faults. This approach requires replicas to agree upon the order of incoming requests and process them in the agreed upon order. Despite numerous efforts to improve the performance or fault scalability of BFT systems [4, 7, 15, 25, 1, 13], existing approaches remain expensive, requiring at least $2f + 1$ replicas to execute each request in order to tolerate f faults [15, 26]. This high replication cost has been a significant barrier to their adoption—to the best of our knowledge, no commercial data center application uses BFT techniques today, despite the wealth of research in this area.

Many recent efforts have focused on optimizing the agreement protocol used by BFT replicas [7, 15]; consequently, today's state-of-the-art protocols can scale to a throughput of 80,000 requests/s and incur overheads of less than 10 μs per request for reaching *agreement* [15]. In contrast, request *execution* overheads for typical applications such as web servers and databases [25] can be in the order of milliseconds or tens of milliseconds—three orders of magnitude higher than the agreement cost. Since request executions dominate the total cost of processing requests in BFT services, the hardware (server) capacity needed for request executions will far exceed that for running the agreement protocol. Hence, we argue that the total cost of a BFT service can be truly reduced only when the total overhead of request executions, rather than the cost to reach agreement, is somehow reduced.

In this paper, we present ZZ, a new approach that reduces the cost of replication as well as that of request executions in BFT systems. Our approach enables general BFT services to be constructed with a replication cost close to $f + 1$, halving the $2f + 1$ or higher cost incurred by state-of-the-art approaches [26]. ZZ targets shared hosting data center environments where replicas from multiple applications can share a physical server. The key insight in ZZ[1] is to run only $f + 1$ execution replicas per application in the graceful case where there are no faults, and to use additional sleeping replicas that get activated only upon failures. By multiplexing fewer replicas onto a given set of shared servers, our approach is able to provide more server capacity to each replica, and thereby achieve higher throughput and lower response times for request executions. In the worst case where all applications experience simultaneous faults, our approach requires an additional f replicas per application, matching the overhead of the $2f + 1$ approach. However, in the common case where only a *subset* of the data center applications are experiencing faults, our approach requires fewer replicas in total, yielding response time and throughput benefits. Like [26], our system still requires $3f + 1$ agreement replicas; however, we argue that the overhead imposed by agreement replicas is small, allowing such replicas from multiple applications to be densely packed onto physical servers.

The ability to quickly activate additional replicas upon fault detection is central to our ZZ approach. While any mechanism that enables fast replica activation can be employed in ZZ, in this paper, we rely upon virtualization—a

[1] Denotes sleeping replicas; from the sleeping connotation of the term "zz.."

	PBFT'99	SEP'03	Zyzzyva'07	ZZ
Agreement replicas	$3f+1$	$3f+1$	$3f+1$	$3f+1$
Execution replicas	$3f+1$	$2f+1$	$2f+1$	$(1+r)f+1$
Agreement MACs/req per replica	$2+\frac{8f+1}{b}$	$2+\frac{12f+3}{b}$	$2+\frac{3f}{b}$	$2+\frac{10f+3}{b}$
Minimum work/requ (for large b)	$(3f+1)\cdot(E+2\mu)$	$(2f+1)E+(3f+1)2\mu$	$(2f+1)E+(3f+1)2\mu$	$(f+1)E+(3f+1)2\mu$
Maximum throughput (if $E \gg \mu$)	$\frac{1}{(3f+1)E}$	$\frac{1}{(2f+1)E}$	$\frac{1}{(2f+1)E}$	$\frac{1}{(f+1)E}$

Table 1. ZZ versus existing BFT approaches. Here, f is the number of allowed faults, b is the batch size, E is execution cost, μ the cost of a MAC operation, and $r \ll 1$ is a variable formally defined in §4.3.3. All numbers are for periods when there are no faults and the network is well-behaved.

technique already employed in modern data centers—for on-demand replica activation.

This paper makes the following contributions. First, we propose a practical solution to reduce the cost of BFT to nearly $f+1$ execution replicas and define formal bounds on ZZ's replication cost. Second, reducing the execution cost in ZZ comes at the expense of potentially allowing faulty nodes to increase response times; we analyze and bound this response time inflation and show that in realistic scenarios malicious applications cannot significantly reduce performance. Finally, we implement a prototype of ZZ by enhancing the BASE library and combining it with the Xen virtual machine and the ZFS file system. ZZ leverages virtualization for fast replica activation and optimizes the recovery protocol to allow newly-activated replicas to immediately begin processing requests through an amortized state transfer strategy. We evaluate our prototype using a BFT web server and a ZZ-based NFS file server. Our experimental results demonstrate that in a prototype data center running four BFT web servers, ZZ's use of only $f+1$ execution replicas in the fault-free case yields response time and throughput improvements of up to 66%, and is still able to rapidly recovery after simultaneous failures occur. Overall, our evaluation emphasizes the importance of minimizing the execution cost of real BFT services and demonstrates how ZZ provides strong fault tolerance guarantees at significantly lower cost compared to existing systems.

2 State-of-the-art vs. the Art of ZZ

In this section, we compare ZZ to state-of-the-art approaches and describe how we reduce the execution cost to $f+1$.

2.1 From 3f+1 to 2f+1

In the traditional PBFT approach [4], during graceful execution, a client sends a request Q to the replicas. The $3f+1$ (or more) replicas agree upon the sequence number corresponding to Q, execute it in that order, and send responses back to the client. When the client receives $f+1$ valid and matching responses from different replicas, it knows that at least one correct replica executed Q in the correct order. Figure 1(a) illustrates how the principle of separating agreement from execution can reduce the number of execution replicas required to tolerate up to f faults from $3f+1$ to $2f+1$. In this separation approach [26], the client sends Q to a primary in the agreement cluster consisting of $3f+1$ lightweight machines that agree upon the sequence number i corresponding to Q and send $[Q,i]$ to the execution cluster consisting of $2f+1$ replicas that store and process application state. When the agreement cluster receives $f+1$ matching responses from the execution cluster, it forwards the response to the client knowing that at least one correct execution replica executed Q in the correct order. For simplicity of exposition, we have omitted cryptographic operations above.

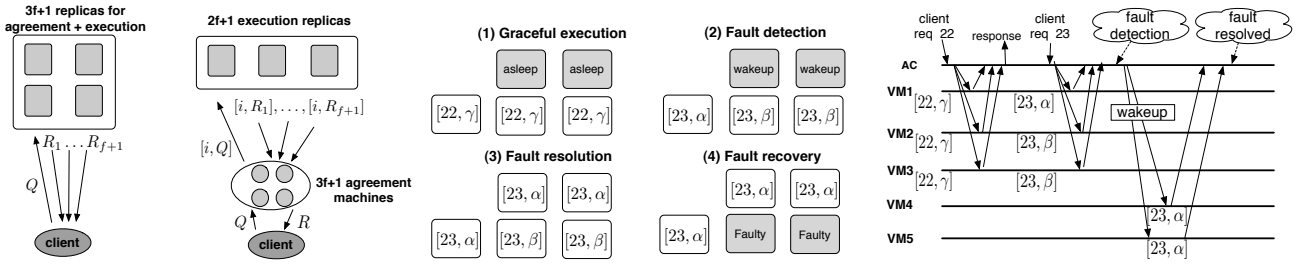
2.2 Circumventing 2f+1

The $2f+1$ replication cost is believed necessary [15, 7, 1] for BFT systems. However, more than a decade ago, Castro and Liskov concluded their original paper on PBFT [4] saying "it is possible to reduce the number of copies of the state to $f+1$ but the details remain to be worked out". In this paper, we work out those details.

Table 1 shows the replication cost and performance characteristics of several BFT State Machine Replication (BFT-SMR) approaches in comparison to ZZ; quorum based approaches such as [7, 1] lead to a similar comparison. All listed numbers are for gracious execution, i.e., when there are no faults and the network is well-behaved. Note that all approaches require at least $3f+1$ replicas in order to tolerate up to f independent Byzantine failures, consistent with classical results that place a lower bound of $3f+1$ replicas for a safe Byzantine consensus protocol that is live under weak synchrony assumptions [10].

In contrast to common practice, we do not measure replication cost in terms of the total number of physical machines as we assume a virtualized environment that is common in many data centers today. Virtualization allows resources to be allocated to a replica at a granularity finer than an entire physical machine. Virtualization itself is useful in multiplexed environments, where a data center owner hosts many services simultaneously for better management of limited available resources. Note that virtualization helps all BFT approaches, not just ZZ, in multiplexed environments.

Cost: Our position is that execution, not agreement, is the dominant provisioning cost for most realistic data center services that can benefit from the high assurance provided by BFT. To put this in perspective, consider that state-of-the-art BFT approaches such as Zyzzyva show a peak throughput of over 80K requests/second for a toy application consisting of *null* requests, which is almost three orders of magnitude more than the achievable throughput for a database service on comparable hardware [25]. Thus in realistic systems, the primary



(a) Separation of Agreement & Execution (b) Protocol for f=2 (c) Fault and Recovery Timeline

Figure 1. (a) The PBFT approach versus the separation of agreement from execution. **(b-c)** Various scenarios in the ZZ system for $f = 2$ faults. Request 22 results in matching responses γ, but the mismatch in request 23 initiates new virtual machine replicas on demand.

cost is that of hardware performing application execution, not agreement. ZZ nearly halves the data center provisioning cost by reducing the number of replicas actively executing requests (Table 1 row 2).

Throughput: ZZ can achieve a higher peak throughput compared to state-of-the-art approaches when execution dominates request processing cost and resources are constrained. For a fair comparison, assume that all approaches are provisioned with the same total amount of resources. Then, the peak throughput of each approach is bounded by the minimum of its best-case execution throughput and its best-case agreement throughput (row 4). Agreement throughput is primarily limited by the overhead μ of a MAC operation and can be improved significantly through batching. However, batching is immaterial to the overall throughput when execution is the bottleneck (row 5).

The comparison above is for performance during periods when there are no faults and the network is well-behaved. In adverse conditions, the throughput and latency of all approaches can degrade significantly and a thorough comparison is nontrivial and difficult to characterize concisely [23, 6].

When failures occur, ZZ incurs a higher latency to execute some requests until its failure recovery protocol is complete. Our experiments suggest that this additional overhead is modest and is small compared to typical WAN delays. In a world where failures are the uncommon case, ZZ offers valuable savings in replication cost or, equivalently, improvement in throughput under limited resources.

ZZ is not a new "BFT protocol" as that term is typically used to refer to the agreement protocol; instead, ZZ is an execution approach that can be interfaced with existing BFT-SMR agreement protocols. Our prototype uses the BASE implementation of the PBFT protocol as it was the most mature and readily available BFT implementation at the time of writing. The choice was also motivated by our premise that we do not seek to optimize agreement throughput, but to demonstrate the feasibility of ZZ's execution approach with a reasonable agreement protocol. Admittedly, it was easier to work out the details of augmenting ZZ to PBFT compared to more sophisticated agreement protocols.

3 ZZ design

3.1 System and Fault Model

We assume a Byzantine failure model where faulty replicas or clients may behave arbitrarily. There are two kinds of replicas: 1) agreement replicas that assign an order to client requests and 2) execution replicas that maintain application state and execute client requests. Replicas fail independently, and we assume an upper bound g on the number of faulty agreement replicas and a bound f on the number of faulty execution replicas in a given window of vulnerability. We initially assume an infinite window of vulnerability, and relax this assumption in Section 4.3.4. An adversary may coordinate the actions of faulty nodes in an arbitrary malicious manner. However, the adversary can not subvert standard cryptographic assumptions about collision-resistant hashes, encryption, and digital signatures.

ZZ uses the state machine replication model to implement a BFT service. Replicas agree on an ordering of incoming requests and each execution replica executes all requests in the same order. Like all previous SMR based BFT systems, we assume that either the service is deterministic or the non-deterministic operations in the service can be transformed to deterministic ones via the agreement protocol [4].

Our system ensures safety in an asynchronous network that can drop, delay, corrupt, or reorder messages. Liveness is guaranteed only during periods of synchrony when there is a finite but possibly unknown bound on message delivery time. The above system model and assumptions are similar to those assumed by many existing BFT systems [4, 15, 26, 22].

Virtualization: ZZ assumes that replicas are being run inside virtual machines. As a result, it is possible to run multiple replicas on a single physical server. To maintain the fault independence requirement, no more than one agreement replica and one execution replica of each service can be hosted on a single physical server.

ZZ assumes that the hypervisor may be Byzantine. Because of the placement assumption above, a malicious hypervisor is equivalent to a single fault in each service hosted on the physical machine. As before, we assume a bound f on the number of faulty hypervisors within a window of vulnerability. We note that even today sufficient hypervisor diversity

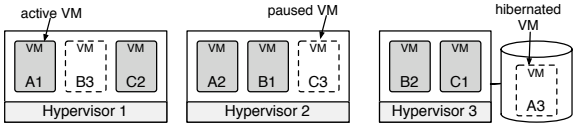


Figure 2. An example server setup with three $f = 1$ fault tolerant applications, A, B, and C; only execution replicas are shown.

(e.g., Xen, KVM, VMWare, Hyper-V) is available to justify this assumption.

3.2 ZZ Design Overview

ZZ reduces the replication cost of BFT from $2f + 1$ to nearly $f + 1$ using virtualization based on two simple insights. First, if a system is designed to be correct in an asynchronous environment, it must be correct even if some replicas are out of date. Second, during fault-free periods, a system designed to be correct despite f Byzantine faults must be unaffected if up to f replicas are turned off. ZZ leverages the second insight to turn off f replicas during fault-free periods requiring just $f+1$ replicas to actively execute requests. When faults occur, ZZ leverages the first insight and behaves exactly as if the f standby replicas were slow but correct replicas.

If the $f + 1$ active execution replicas return matching responses for an ordered request, at least one of these responses, and by implication all of the responses, must be correct. The problematic case is when the $f + 1$ responses do not match. In this case, ZZ starts up additional virtual machines hosting standby replicas For example, when $f = 1$, upon detecting a fault, ZZ starts up a third replica that executes the most recent request. Since at most one replica can be faulty, the third response must match one of the other two responses, and ZZ returns this matching response to the client. Figure 1(b-c) illustrates the high-level control flow for $f = 2$. Request 22 is executed successfully generating the response γ, but request 23 results in a mismatch waking up the two standby VM replicas. The fault is resolved by comparing the outputs of all $2f + 1$ replicas, revealing α as the correct response.

The above design would be impractical without a quick replica wake-up mechanism. Virtualization provides this mechanism by maintaining additional replicas in a "dormant" state. Figure 2 illustrates how ZZ can store additional replicas both in memory as prespawned but paused VMs and hibernated to disk. Paused VMs resume within milliseconds but consume memory resources. Hibernated replicas require only storage resources, but can incur greater recovery times.

3.3 Design Challenges

The high-level approach described above raises several further challenges. First, how does a restored replica obtain the necessary application state required to execute the current request? In traditional BFT systems, each replica maintains an independent copy of the entire application state. Periodically, all replicas create application checkpoints that can be used to bring up to speed any replicas which become out of date. However, a restored ZZ replica may not have any previous version of application state. It must be able to verify that the state it obtains is correct even though there may be only one correct execution replica (and f faulty ones), e.g., when $f = 1$, the third replica must be able to determine which of the two existing replicas possesses the correct state.

Second, transferring the entire application state can take an unacceptably long time. In existing BFT systems, a recovering replica may generate incorrect messages until it obtains a stable checkpoint. This inconsistent behavior during checkpoint transfer is treated like a fault and does not impede progress of request execution if there is a quorum of $f + 1$ correct execution replicas with a current copy of the application state. However, when a ZZ replica recovers, there may exist just one correct execution replica with a current copy of the application state. The traditional state transfer approach can stall request execution in ZZ until f recovering replicas have obtained a stable checkpoint.

Third, ZZ's replication cost must be robust to faulty replica or client behavior. A faulty client must not be able to trigger recovery of standby replicas. A compromised replica must not be able to trigger additional recoveries if there are at least $f + 1$ correct and active replicas. If these conditions are not met, the replication cost savings would vanish and system performance could be worse than a traditional BFT system using $2f + 1$ replicas.

4 ZZ Protocol

In this section we briefly describe the separated protocol from [26], and present ZZ's modifications to support switching from $f + 1$ to $2f + 1$ execution replicas after faults are detected.

4.1 Graceful Execution

Client Request & Agreement: In Figure 3 step 1, a client c sends a request Q to the agreement cluster to submit an operation o with a timestamp t. The timestamps ensure exactly-once semantics for execution of client requests, and a faulty client's behavior does not affect other clients' requests.

Upon receiving a client request Q, the agreement replicas will execute the standard three phase BFT agreement protocol [4] in order to assign a sequence number n to the request. When an agreement replica j learns of the sequence number n committed to Q, it sends a commit message C to all execution replicas (Fig. 3 step 2).

Execution: An execution replica i executes a request Q when it gathers a commit certificate $\{C_i\}, i \in A|2g + 1$—a set of $2g + 1$ valid and matching commit messages from the agreement cluster, and it has executed all other requests with a lower sequence number. Each execution node produces a reply R which it sends to the client and an execution report message, *ER*, which is sent to all agreement nodes (Fig. 3 steps 3 and 4).

In the normal case, the client receives a response certificate $\{R_i\}, i \in E|f+1$—matching reply messages from $f+1$ execution nodes. Since at most f execution replicas can be faulty, a client receiving a response certificate knows that the response is correct. If a client does not obtain matching replies, it resends its request to the agreement cluster. If an agreement node receives a retransmitted request for which it has received $f+1$ matching execution report messages, then it can send a reply affirmation, *RA* to the client (Fig. 3 step 5). If a client receives $g+1$ such messages containing a response digest, $\overline{R}$, matching one of the replies already received, then the client can accept that reply as valid. This "backup" solution is used by ZZ to prevent unnecessary wakeups where a partially faulty execution node may reply to the agreement cluster, but not to the client. If the agreement cluster cannot produce an affirmation for the client, then additional nodes must be started as described in subsequent sections.

4.2 Dealing with Faults

4.2.1 Checkpointing

Checkpoints are used so that newly started execution replicas can obtain a recent copy of the application state and so that replicas can periodically garbage collect their logs. The checkpoints are constructed at predetermined request sequence numbers, e.g., when it is exactly divisible by 1024.

With at least $2f+1$ execution replicas in other BFT-SMR systems, a recovering replica is guaranteed to get at least $f+1$ valid and matching checkpoint messages from other execution replicas, allowing checkpoint creation and validation to be done exclusively within the execution cluster [26]. However, ZZ runs only $f+1$ execution replicas in the normal case, and thus a new replica may not be able to tell which of the checkpoints it is provided are correct.

To address this problem, ZZ's execution cluster must coordinate with the agreement cluster during checkpoint creation. ZZ execution nodes create checkpoints of application state and their reply log, then assemble a proof of their checkpoint, *CP*, and send it to all of the execution *and* agreement nodes (Fig. 3 step 6). Informing the agreement nodes of the checkpoint digest allows them to assist recovering execution replicas in verifying the checkpoint data they obtain from potentially faulty nodes.

A checkpoint certificate $\{CP_i\}, i \in E|f+1$ is a set of $f+1$ *CP* messages with matching digests. When an execution replica receives a checkpoint certificate with a sequence number n, it considers the checkpoint as stable and discards earlier checkpoints and request commit certificates with lower sequence numbers that it received from the agreement cluster. Likewise, the agreement nodes are able to use these checkpoint certificates to determine when they can garbage collect messages in their communication log with the execution cluster.

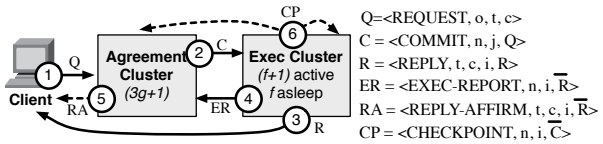


Figure 3. The normal agreement and execution protocol in ZZ proceeds through steps 1-4. Step 5 is needed only after a fault. Checkpoints (step 6) are created on a periodic basis. The notation $\langle \text{LABEL}, X \rangle$ denotes the message of type LABEL with parameters X. We indicate the digest of parameter Y as $\overline{Y}$.

4.2.2 Fault Detection

The agreement cluster is responsible for detecting faults in the execution cluster. Agreement nodes in ZZ are capable of detecting invalid execution or checkpoint messages; the fault detection and recovery steps for each of these are identical, so for brevity we focus on invalid or missing execution responses. In the normal case, an agreement replica j waits for an execution certificate, $\{ER_i\}, i \in E|f+1$, from the execution cluster. Replica j inserts this certificate into a local log ordered by the sequence number of requests. When j receives *ER* messages which do not match, or waits for longer than a predetermined timeout, j sends a *recovery request*, $W = \langle \text{RECOVER}, j, n \rangle_j$, to the f hypervisors controlling the standby execution replicas. When the hypervisor of a sleeping execution replica receives a recovery certificate, $\{W_i\}, i \in A|g+1$, it wakes up the local execution replica.

4.2.3 Replica Recovery with Amortized State Transfer

When an execution replica k starts up, it must obtain the most recent checkpoint of the entire application state from existing replicas and verify that it is correct. Unfortunately, checkpoint transfer and verification can take an unacceptably long time. Worse, unlike previous BFT systems that can leverage copy-on-write techniques and incremental cryptography schemes to transfer only the objects modified since the last checkpoint, a recovering ZZ replica has no previous checkpoints.

How does replica k begin to execute requests without any application state? Instead of performing an expensive transfer of the entire state upfront, a recovering ZZ replica fetches and verifies the state necessary to execute each request on demand. Replica k first fetches a log of committed requests since the last checkpoint from the agreement cluster and a checkpoint certificate $\{CP_i\}, i \in E, A|g+1$ from some combination of $g+1$ execution and agreement replicas. This checkpoint certificate includes digests for each state object, allowing the replica to verify that any state object it obtains has come from a correct replica.

After obtaining the checkpoint certificate with object digests, replica k begins to execute in order the recently committed requests. Let Q be the first request that reads from or writes to some object p since the most recent checkpoint. To execute Q, replica k fetches p on demand from any execution replica that can provide an object consistent with p's digest that k learned from the certificate. Replica k continues

executing requests in sequence number order fetching new objects on demand until it obtains a stable checkpoint.

Recovery is complete only when replica k has obtained a stable checkpoint. Since on-demand fetches only fetch objects touched by requests, they are not sufficient for k to obtain a stable checkpoint, so the replica must also fetch the remaining state in the background.

4.3 System Properties

Here we formally define the performance, replication cost, safety and liveness properties of ZZ. Due to space constraints we defer complete proofs to [3].

4.3.1 Response Time Inflation

ZZ relies on timeouts to detect faults in execution replicas. This opens up a potential performance vulnerability. A low value of the timeout can trigger fault detection even when the delays are benign and needlessly start new replicas. On the other hand, a high value of the timeout can be exploited by faulty replicas to degrade performance as they can delay sending each response to the agreement cluster until just before the timeout. The former can take away ZZ's savings in replication cost as it can end up running more than $f+1$ (and up to $2f+1$) replicas even during graceful periods. The latter hurts performance under faults. Note that safety is not violated in either case.

To address this problem, we suggest the following simple procedure for setting timeouts to limit response time inflation. Upon receiving the first response to a request committed to sequence number n, an agreement replica sets the timeout τ_n to Kt_1, where t_1 is the response time of the first response and K is a pre-configured variance bound. If the agreement replica does not receive f more matching responses within τ_n, then it triggers a fault and wakes up f additional replicas.

This procedure trivially bounds the response time inflation of requests to a factor of K, but we can further constrain the performance impact by considering the response time distribution as follows. Given p, the probability of a replica being faulty,

THEOREM 1 *Faulty replicas can inflate average response time by a factor of* $min(1, I)$*:*

$$I = \sum_{1 \le m \le f} p^m \frac{K \cdot E[\mathit{MIN}_{f+1-m}]}{E[\mathit{MAX}_{f+1}]}$$

where $E[\mathit{MIN}_{f+1-m}]$ is the expected minimum response time for a set of $f+1-m$ replicas and $E[\mathit{MAX}_{f+1}]$ is the expected maximum response time of all $f+1$ replicas, assuming all response times are identically distributed as some distribution Ψ. The top term in this expression follows from the rule defined above: a faulty node can increase response time by at most K compared to the fastest correct replica (i.e. the replica with the minimum response time out of $f+1-m$ nodes). The bottom term is the non-faulty case where response time is limited by the slowest of the $f+1$ replicas. We must sum this fraction over all possible cases, $m = 1, 2, ...f$ faults. As an example, suppose $K = 4$, $f = 3$, and response times are exponentially distributed with $E[\Psi] = 2ms$. Then $E[\mathit{MIN}_{f+1-m}] = \frac{2}{3+1-m}ms$ and $E[\mathit{MAX}_{f+1}] = 4.2ms$. If $p = 0.1$, then $I = 0.075$, i.e. no inflation of average response time. Only for $p > 0.56$ is $I > 1$; if $p = 0.75$ then $I = 1.81$. Note that proactive recovery can be used to ensure p remains small [5] and that to achieve this worst case bound faulty nodes must be able to predict the earliest response time of correct replicas. In practice, correct execution replicas may sometimes violate the variance bound due to benign execution or network delays, causing a *false timeout*. These false timeouts can impact overall replication cost as described in section 4.3.3.

4.3.2 Waking Up and Shutting Down

Since waking up nodes to respond to faults is an expensive procedure, ZZ distinguishes between "blocking" and "non-blocking" faults, and only triggers a wakeup event for blocking faults—those which cannot be resolved without a wakeup. Fortunately, blocking faults by definition are more widespread in their impact, and thus can always be traced back to a faulty execution node which can then be shutdown.

THEOREM 2 *If a wakeup occurs, ZZ will be able to terminate at least one faulty or slow replica.*

To understand the difference, consider the response matrix where position (i, j) indicates E_i's response as reported by agreement node A_j. Consider two examples where the client receives conflicting responses P and Q, and $f = g = 1$,

Non-blocking fault	**Blocking fault**
$A_1 A_2 A_3 A_4$	$A_1 A_2 A_3 A_4$
$E_1 : Q\ P\ P\ P$	$E_1 : Q\ P\ P\ P$
$E_2 : Q\ P\ P\ P$	$E_2 : Q\ Q\ Q\ P$

In the first scenario, it is impossible to distinguish whether only A_1 is faulty or if an execution replica and A_1 is faulty; however, $g+1$ agreement nodes can provide a reply affirmation that P is the correct response. In the second case, there is no way to tell whether Q or P is the correct response, so a wakeup is required. Once this replica is started, it will be simple to determine whether E_1 or E_2 should be terminated.

4.3.3 Overall Replication Cost

The expected replication cost of ZZ varies from $f+1$ to $2f+1$ depending on the probability of replicas being faulty p, and the likelihood of false timeouts, Π_1.

THEOREM 3 *The expected replication cost of ZZ is less than* $(1+r)f+1$*, where* $r = 1-(1-p)^{f+1} + (1-p)^{f+1}\Pi_1$.

These two factors influence the replication cost because additional nodes are started only if 1) a replica truly is faulty (which happens with probability $1-(1-p)^{f+1}$), or 2) there are no faults, but a correct slow replica causes a false timeout (which happens with probability $(1-p)^{f+1}\Pi_1$). In either

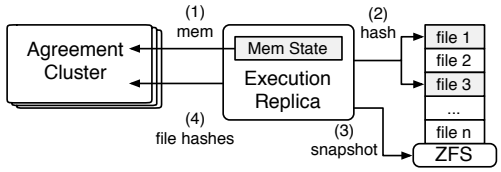


Figure 4. For each checkpoint an execution replica (1) sends any modified memory state, (2) creates hashes for any modified disk files, (3) creates a ZFS snapshot, and (4) returns the list of hashes to agreement nodes.

case, the replication cost is increased by f, resulting in the theorem. The value of p can be reduced by proactive recovery, and Π_1 is dependent on the value of K. Adjusting K results in a tradeoff between the replication cost and the response time inflation bound. Note that in practice the replication cost may be even lower than this because ZZ will quickly shutdown nodes after the fault has been resolved.

4.3.4 Safety and Liveness Properties

ZZ provides the following safety and liveness properties. The proofs for these properties follow from those of Separated [26] and PBFT [5].

ZZ ensures the safety property that if a correct client obtains either a response certificate or an affirmation certificate for a response $\langle \text{REPLY}, t, c, j, R \rangle_j$, then (1) the client issued a request $\langle \text{REQUEST}, o, t, c \rangle_c$ earlier; (2) all correct replicas agree on the sequence number n of that request and on the order of all requests with sequence numbers in $[1, n]$; (3) the value of the reply R is the reply that a single correct replica would have produced if it started with a default initial state S_0 and executed each operation $o_i, 1 \leq i \leq n$, in that order, where o_i denotes the operation requested by the request to which sequence number i was assigned.

ZZ also ensures the liveness property that if a correct client sends a request R with a timestamp exceeding previous requests and repeatedly retransmits the request, then it will eventually receive a response certificate or an affirmation certificate for R.

Reducing the Window of Vulnerability: ZZ's current implementation assumes an infinite window of vulnerability, i.e., the length of time in which up to f faults can occur. However, this assumption can be relaxed using proactive recovery [5]. By periodically forcing replicas to recover to a known clean state, proactive recovery allows for a configurable finite window of vulnerability.

5 ZZ Implementation

We implemented ZZ by enhancing the 2007 version of BASE [22] so as to 1) use virtual machines to run replicas, 2) incorporate ZZ's checkpointing, fault detection, rapid recovery and fault-mode execution mechanisms, and 3) use file system snapshots to assist checkpointing.

5.1 Replica Control Daemon

We have implemented a ZZ replica control daemon that runs on each physical machine and is responsible for managing replicas after faults occur. The control daemon, which runs in Xen's Domain-0, uses the certificate scheme described in Section 4.2.2 to ensure that it only starts or stops replicas when enough non-faulty replicas agree that it should do so.

Inactive replicas are maintained in either a paused state, where they have no CPU cost but incur a small memory overhead on the system, or hibernated to disk which utilizes no resources other than disk space. To optimize the wakeup latency of replicas hibernating on disk, ZZ uses a paged-out restore technique that exploits the fact that hibernating replicas initially have no useful application state in memory, and thus can be created with a bare minimum allocation of 128MB of RAM (which reduces their disk footprint and load times). After being restored, their memory allocation is increased to the desired level. Although the VM will immediately have access to its expanded memory allocation, there may be an application dependent period of reduced performance if data needs to be paged in.

5.2 Exploiting File System Snapshots

Checkpointing in ZZ relies on the existing mechanisms in the BASE library to save the protocol state of the agreement nodes and any memory state used by the application on the execution nodes. In addition, ZZ supports using the snapshot mechanism provided by modern journaled file systems [27] to simplify checkpointing disk state. Creating disk snapshots is efficient because copy-on-write techniques prevent the need for duplicate disk blocks to be created, and the snapshot overhead is independent of the disk state of the application. ZZ uses ZFS for snapshot support, and works with both the native Solaris and user-space Linux ZFS implementations.

ZZ includes meta-information about the disk state in the checkpoint so that the recovery nodes can validate the disk snapshots created by other execution nodes. To do so, execution replicas create a cryptographic hash for each file in the disk snapshot and send it to the agreement cluster as part of the checkpoint certificate as shown in Figure 4. Hashes are computed only for those files that have been modified since the previous epoch; hashes from the previous epoch are reused for unmodified files to save computation overheads.

Tracking Disk State Changes: The BASE library requires all state, either objects in memory or files on disk, to be registered with the library. In ZZ we have simplified the tracking of disk state so that it can be handled transparently without modifications to the application. We define functions bft_fopen() and bft_fwrite() which replace the ordinary fopen() and fwrite() calls in an application. The bft_fwrite() function invokes the modify() call of the BASE library which must be issued whenever a state object is being edited. This ensures that any files which are modified during an epoch will be rehashed during checkpoint creation.

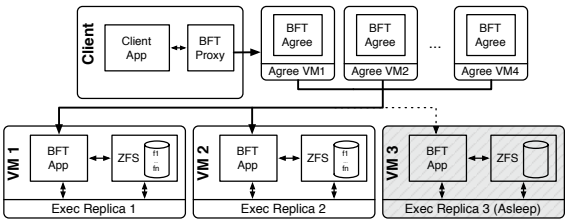


Figure 5: Experimental setup for a basic ZZ BFT service.

For the initial execution replicas, the bft_fopen() call is identical to fopen(). However, for the additional replicas which are spawned after a fault, the bft_fopen call is used to retrieve a file from the disk snapshots and copy it to the replica's own disk on demand. When a recovering replica first tries to open a file, it calls bft_fopen(foo), but the replica will not yet have a local copy of the file. The recovery replica fetches a copy of the file from any replica and verifies it against the hash contained in the most recent checkpoint. If the hashes do not match, the recovery replica requests the file from a different replica, until a matching copy is found and copied to its own disk.

6 Experimental Evaluation

6.1 Experiment Setup

Our experimental data-center setup uses a cluster of 2.12 GHz 64-bit dual-core Dell servers, each with 4GB RAM. Each machine runs a Xen v3.1 hypervisor and Xen virtual machines. Both domain-0 (the controller domain in Xen) as well as the individual VMs run the CentOS 5.1 Linux distribution with the 2.6.18 Linux kernel and the user space ZFS filesystem. All machines are interconnected over gigabit ethernet. Figure 5 shows the setup for agreement and execution replicas of a generic BFT app for $g = f = 1$; multiple such applications are assumed to be run in a BFT data center.

Our experiments involve three fault-tolerant server applications: a Web Server, an NFS server, and a toy client-server microbenchmark.

Fault-tolerant Web Server: We have implemented a BFT-aware HTTP 1.0 Web server that mimics a dynamic web site with server side scripting. The request execution time is configurable to simulate more complex request processing. We generate web workloads using *httperf* clients which contact a local BFT web proxy that submits the requests to the agreement nodes.

Fault-tolerant NFS: BASE provides an NFS client relay and a BFT wrapper for the standard NFS server. We have extended this to support ZZ's on demand state transfer which allows a recovery replica to obtain file system state from ZFS snapshots as it processes each request.

Client-Server Microbenchmark: We utilize the simple client-server application from the BASE library to measure ZZ's performance for *null* requests and to study it's recovery costs under different application state scenarios.

	Graceful performance				After failure		
	h_1	h_2	h_3	h_4	h_1	h_2	h_3
BASE	1234	1234	1234	1234	1234	1234	1234
SEP_{Agree}	1234	1234	1234	1234	1234	1234	1234
SEP_{Exec}	134	124	123	234	134	124	123
ZZ_{Agree}	1234	1234	1234	1234	1234	1234	1234
ZZ_{Exec}	12	12	34	34	12**3**	12**4**	34
ZZ_{Sleep}	3	4	1	2			1

Table 2. Placement of the 4 web servers' virtual machines (denoted 1 to 4) on the 4 data center hosts (h_1 to h_4) under graceful performance and on the 3 remaining hosts after h_4 failure.

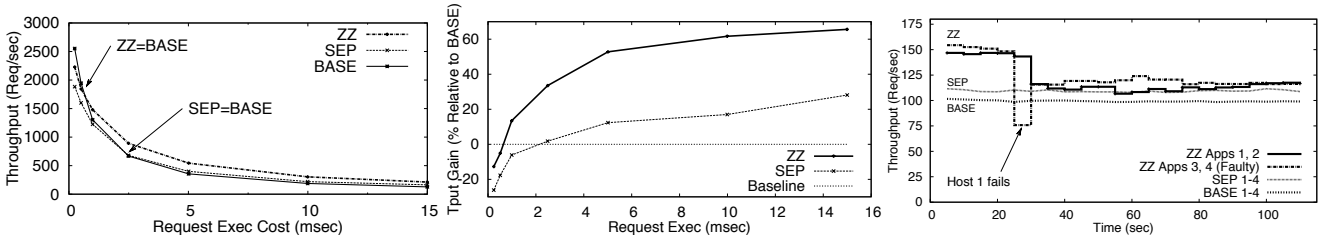
Our experiments compare three systems: ZZ, BASE, and Separated (SEP). BASE is the standard BFT library described in [22]. SEP is our extension of BASE which separates the agreement and execution replicas, and requires $3f + 1$ agreement and $2f + 1$ execution replicas similar to [26]. ZZ also requires $3f + 1$ agreement replicas, but extends SEP to use only $f + 1$ active execution replicas, with an additional f sleeping replicas. While more recent agreement protocols provide higher performance than BASE, our evaluation focuses on cases where execution is at least an order of magnitude more expensive than agreement; we believe our conclusions are consistent with what would be found with more optimized agreement protocols.

6.2 Graceful Performance

We study the graceful performance of ZZ by emulating a shared hosting environment running four independent web apps on four machines. Table 2 shows the placement of agreement and execution replicas on the four hosts. As the agreement and execution clusters can independently handle faults, each host can have at most one replica of each type per application.

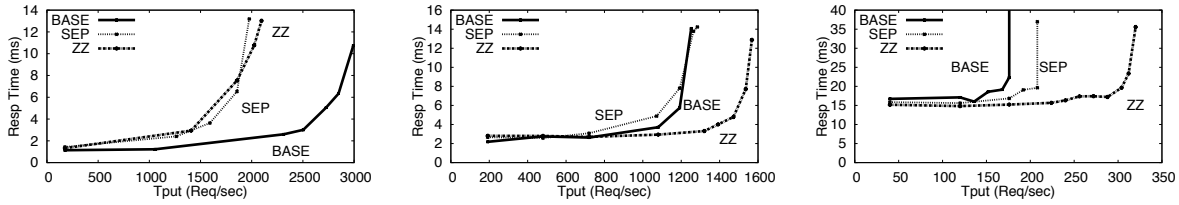
6.2.1 Throughput

We first analyze the impact of request execution cost under ZZ, SEP, and BASE, which require $f+1$, $2f+1$, and $3f+1$ execution replicas per web server respectively. Figure 6(a) compares the throughput of each system as the execution cost per web request is adjusted. When execution cost averages 100 μs, BASE performs the best since the agreement overhead dominates the cost of processing each request and our implementation of separation incurs additional cost for the agreement replicas. However, for execution costs exceeding 0.75 ms, the execution replicas become the system bottleneck. As shown in Figure 6(b), ZZ begins to outperform BASE at this point, and performs increasingly better compared to both BASE and SEP as execution cost rises. SEP surpasses BASE for request costs over 2ms, but cannot obtain the throughput of ZZ since it requires $2f + 1$ replicas instead of only $f+1$. ZZ provides as much as a 66% increase in application throughput relative to BASE for requests with large execution costs.



(a) Application Throughput (b) ZZ Throughput Gain (c) Simultaneous Failures

Figure 6. (a-b) When resources are constrained, ZZ significantly increases system throughput by using fewer replicas. **(c)** Under simultaneous failures of several applications, ZZ quickly recovers and still maintains good throughput.



(a) Null Request (b) 1 ms Request (c) 10 ms Request

Figure 7: For high execution costs, ZZ achieves both higher throughput and lower response times.

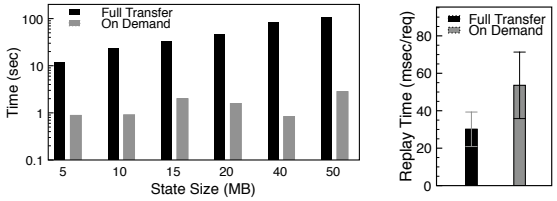
6.2.2 Latency

This experiment further characterizes the performance of ZZ in graceful operation by examining the relation between throughput and response time for different request types. Figure 7 shows the relation between throughput and response time for increasingly CPU intensive request types. For null requests or at very low loads, Figure 7(a), BASE beats SEP and ZZ because it has less agreement overhead. At 1ms, ZZ's use of fewer execution replicas enables it to increase the maximum throughput by 25% over both SEP and BASE. When the execution cost reaches 10ms, SEP outperforms BASE since it uses $2f + 1$ instead of $3f + 1$ replicas. ZZ provides a 33% improvement over SEP, showing the benefit of further reducing to $f + 1$.

6.3 Simultaneous Failures

When several applications are multiplexed on a single physical host, a faulty node can impact all its running applications. In this experiment, we simulate a malicious hypervisor on one of the four hosts that causes multiple applications to experience faults simultaneously. Host h_4 in Table 2 is set as a faulty machine and is configured to cause faults on all of its replicas 20 seconds into the experiment as shown in Figure 6(c). For ZZ, the failure of h_4 directly impacts web servers 3 and 4 which have active execution replicas there. The replica for server 2 is a sleeping replica, so its corruption has no effect on the system. The failure also brings down one agreement replica for each of the web servers, however they are able to mask these failures since $2f+1$ correct agreement replicas remain on other nodes.

ZZ recognizes the corrupt execution replicas when it detects disagreement on the request output of each service. It responds by waking up the sleeping replicas on hosts h_1 and h_2. After a short recovery period (further analyzed in the next section), ZZ's performance is similar to that of SEP with three active execution replicas competing for resources on h_1 and h_2. Even though h_3 only has two active VMs and uses less resources with ZZ, applications 3 and 4 have to wait for responses from h_1 and h_2 to make progress. Both ZZ and SEP maintain a higher throughput than BASE that runs all applications on all hosts.



(a) Loading Checkpoints (b) Request Replay

Figure 8. (a) The cost of full state transfer increases with state size. (b) On Demand incurs overhead when replaying requests since state objects must be verified.

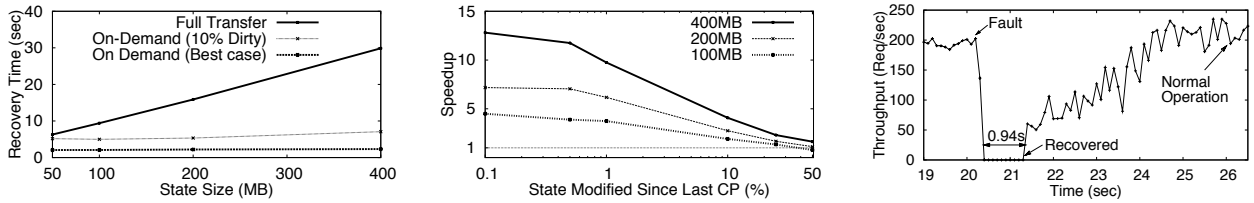
6.4 Recovery Cost

The following experiments study the cost of recovering replicas in more detail using both microbenchmarks and our fault tolerant NFS server. We study the recovery cost, which we define as the delay from when the agreement cluster detects a fault until the client receives the correct response.

6.4.1 NFS Recovery Costs

We investigate the NFS server recovery cost for a workload that creates 200 files of equal size before encountering a fault. We vary the size of the files to adjust the total state maintained by the application, which also impacts the number of requests which need to be replayed after the fault.

ZZ uses an on-demand transfer scheme for delivering ap-



(a) Recovery Time (b) On-Demand Speedup Factor (c) Fault Mode Performance

Figure 9. (a-b) The worst case recovery time depends on the amount of state updated between the last checkpoint and the fault. **(c)** The recovery period lasts for less than a second. At first, requests see higher latency since state must be fetched on-demand.

plication state to newly recovered replicas. Figure 8(a) shows the time for processing the checkpoints when using full transfer or ZZ's on-demand approach (note the log scale). The full state transfer approach performs very poorly since the BFT NFS wrapper must both retrieve the full contents of each file and perform RPC calls to write out all of the files to the actual NFS server. When transferring the full checkpoint, recovery time increases exponentially and state sizes greater than a mere 20 megabytes can take longer than 60 seconds, after which point NFS requests typically will time out. In contrast, the on-demand approach has a constant overhead with an average of 1.4 seconds. This emphasizes the importance of using the on-demand transfer for realistic applications where it is necessary to make some progress in order to prevent application timeouts.

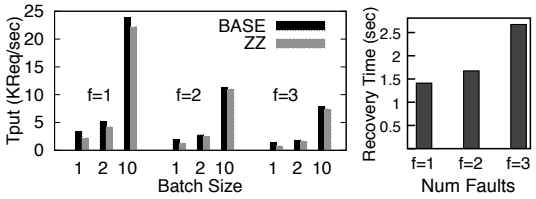
We report the average time per request replayed and the standard deviation for each scheme in Figure 8(b). ZZ's on demand system experiences a higher replay cost due to the added overhead of fetching and verifying state on-demand; it also has a higher variance since the first access to a file incurs more overhead than subsequent calls. While ZZ's replay time is larger, the total recovery time is much smaller when using on-demand transfer.

6.4.2 Obtaining State On-Demand

This experiment uses a BFT client-server microbenchmark which processes requests with negligible execution cost to study the recovery cost after faults are caused in applications with different state sizes.

In the best case, a fault occurs immediately after a checkpoint and new replicas only need to load and resume from the last save, taking a constant time of about 2s regardless of state size (Figure 9(a)). Otherwise, the cost of on-demand recovery varies depending on the amount of application state that was modified since the last checkpoint. The "10% Dirty" line shows the recovery cost when 10% of the application's state needs to be fetched during replay. In that case, ZZ's recovery time varies from 5.2s to 7.1s for states of 50 and 400MB, respectively. This remains much faster than the Full Transfer technique which requires over 30s to transfer and verify 400MB of state.

The tradeoff between amount of dirty state and recovery speed is further studied in Figure 9(b). Even when 10% of application state is modified between each checkpoint, on-demand transfers speed up recovery by at least five times.



(a) Null Request Performance (b) Recovery Time

Figure 10. Recovery time increases for larger f from message overhead and increased ZFS operations.

Only when more than 50% of state is dirty does it becomes more expensive to replay than to perform a full transfer. Fortunately, we have measured the additional cost of ZFS checkpoints at 0.03s, making it practical to checkpoint every few seconds, during which time most applications will only modify a small fraction of their total application state.

Next we examine the impact of on-demand recovery on throughput and latency. The client sends a series of requests involving random accesses to 100KB state objects and a fault is injected after 20.2s (Figure 9(c)). The faulty request experiences a sub-second recovery period, after which the application can handle new requests. The mean request latency prior to the fault is 5ms with very little variation. The latency of requests after the fault has a bimodal distribution depending on whether the request accesses a file that has already been fetched or one which needs to be fetched and verified. The long requests, which include state verification and transfer, take an average of 20ms. As the recovery replica rebuilds its local state, the throughput rises since the proportion of slow requests decreases. After 26s, the full application state has been loaded by the recovery replica, and the throughput prior to the fault is once again maintained.

6.5 Trade-offs and Discussion

6.5.1 Impact of Multiple Faults

We examine how ZZ's graceful performance and recovery time changes as we adjust f, the number of faults supported by the system when *null* requests are used requiring no execution cost. Figure 10(a) shows that ZZ's graceful mode performance scales similarly to BASE as the number of faults increases. This is expected because the number of cryptographic and network operations rises similarly in each.

We next examine the recovery latency of the client-server

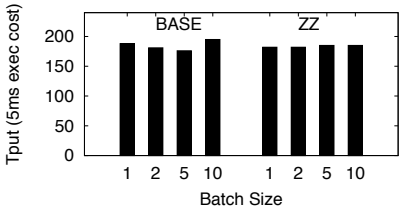


Figure 11. Throughput of ZZ and BASE with different batch sizes for a 5ms request execution time.

microbenchmark for up to three faults. We inject a fault to f of the active execution replicas and measure the recovery time for f new replicas to handle the faulty request. Figure 10(b) shows how the recovery time increases slightly due to increased message passing and because each ZFS system needs to export snapshots to a larger number of recovering replicas. We believe the overhead can be attributed to our use of the user-space ZFS code that is less optimized than the Solaris kernel module implementation, and messaging overhead which could be decreased with hardware multicast.

6.5.2 Agreement Protocol Performance

Various agreement protocol optimizations exist such as request batching, but these may have less effect when request execution costs are non-trivial. While Figure 10(a) shows a large benefit of batching for null requests, Figure 11 depicts a similar experiment with a request execution time of 5ms. We observe that batching improvements become insignificant with non-trivial execution costs. This demonstrates the importance of reducing execution costs, not just agreement overhead, for real applications.

6.5.3 Maintaining Spare VMs

In our previous experiments recovery VMs were kept in a paused state which provides a very fast recovery but consumes memory. Applications that have less stringent latency requirements can keep their recovery VMs hibernated on disk, removing the memory pressure on the system.

With a naive approach, maintaining VMs hibernated to disk can increase recovery latency by a factor proportional to their amount of RAM. This is because restoring a hibernated VM involves loading the VM's full memory contents from disk. The table below shows how our paged-out restore technique reduces the startup time for a VM with a 2GB memory allocation from over 40 seconds to less than 6 seconds.

Operation	Time (sec)
Xen Restore (2GB image)	44.0
Paged-out Restore (128MB→2GB)	5.88
Unpause VM	0.29
ZFS Clone	0.60

ZZ utilizes ZFS to simplify checkpoint creation at low cost. The ZFS clone operation is used during recovery to make snapshots from the previous checkpoint available to the recovery VMs. This can be done in parallel with initializing the recovery VMs, and incurs only minimal latency.

7 Related Work

[19] introduced the problem of Byzantine agreement. Lamport also introduced the state machine replication approach [16] that relies on consensus to establish an order on requests. Consensus in the presence of asynchrony and faults has seen almost three decades of research. [10] established a lower bound of $3f + 1$ replicas for Byzantine agreement given partial synchrony, i.e., an unknown but fixed upper bound on message delivery time. The classic FLP [11] result showed that no agreement protocol is guaranteed to terminate with even one (benignly) faulty node in an asynchronous environment. Viewstamped replication [20] and Paxos [17] describe an asynchronous state machine replication approach that is safe despite crash failures.

Early BFT systems [21, 14] incurred a prohibitively high overhead and relied on failure detectors to exclude faulty replicas. However, accurate failure detectors are not achievable under asynchrony, thus these systems effectively relied on synchrony for safety. Castro and Liskov's PBFT [4] introduced a BFT SMR-based system that relied on synchrony only for liveness. The view change protocol at the core of PBFT shares similarities with viewstamped replication [20] or Paxos [17] but incurs a replication cost of at least $3f + 1$ for safety. PBFT showed that the latency and throughput overhead of BFT can be low enough to be practical. The FARSITE system [2] reduces the replication cost of a BFT file-system to $f + 1$; in comparison, ZZ has similar goals, but is able to provide the same cost reduction for any application which can be represented by a more general SMR system. ZZ draws inspiration from Cheap Paxos [18], which advocated the use of cheaper auxiliary nodes used only to handle crash failures of main nodes. Our contribution lies in extending the idea to Byzantine faults and demonstrating its practicality through a system design and implementation.

Virtualization has been used in several BFT systems recently since it provides a clean way to isolate services. The VM-FIT systems exploits virtualization for isolation and to allow for more efficient proactive recovery [9]. Like ZZ, VM-FIT employs an amortized state transfer mechanism to efficiently update replicas, but it is designed for a system running 2f+1 execution nodes. The idea of "reactive recovery", where faulty replicas are replaced after fault detection, was used in [24], which employed virtualization to provide isolation between different types of replicas. In ZZ, reactive recovery is not an optional optimization, but a requirement since in order to make progress it must immediately instantiate new replicas after faults are detected.

The Remus system uses virtualization to provide black-box crash fault tolerance using a standby VM replica [8]. ZZ seeks to provide stronger Byzantine fault tolerance guarantees at a similar replication cost, although ZZ, like all BFT

systems, requires application support for the BFT protocol. Terra is a virtual machine platform for trusted computing that employs a trusted hypervisor [12]; ZZ allows hypervisors to be Byzantine faulty.

8 Conclusions

In this paper, we presented ZZ a new execution approach that can be interfaced with existing BFT-SMR agreement protocols to reduce the replication cost from $2f + 1$ to practically $f + 1$. Our key insight was to use $f + 1$ execution replicas in the normal case and to activate additional VM replicas only upon failures. We implemented ZZ using the BASE library and the Xen virtual machine and evaluated it on a prototype data center that emulates a shared hosting environment. The key results from our evaluation are as follows. (1) In a prototype data center with four BFT web servers, ZZ lowers response times and improves throughput by up to 66% and 33% in the fault-free case, when compared to systems using $3f + 1$ and $2f + 1$ replicas, respectively. (2) In the presence of multiple application failures, after a short recovery period, ZZ performs as well or better than $2f + 1$ replication and still outperforms BASE's $3f + 1$ replication. (3) The use of paused virtual machine replicas and on-demand state fetching allows ZZ to achieve sub-second recovery times. (4) We find that batching in the agreement nodes, which significantly improves the performance of null execution requests, yields no perceptible improvements for realistic applications with non-trivial request execution costs. Overall, our results demonstrate that, in shared data centers that host multiple applications with substantial request execution costs, ZZ can be a practical and cost-effective approach for providing BFT.

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9 Appendix

9.1 Limiting Response Time Inflation

ZZ triggers a fault if replicas respond more than K times slower than the fastest replica. This procedure trivially bounds the response time inflation of requests to a factor of K, but we can further constrain the performance impact by considering the response time distribution as follows. Given p, the probability of a replica being faulty,

THEOREM 1 *Faulty replicas can inflate average response time by a factor of* $min(1, I)$*:*

$$I = \sum_{1 \leq m \leq f} p^m \frac{K \cdot E[MIN_{f+1-m}]}{E[MAX_{f+1}]}$$

where $E[MIN_{f+1-m}]$ is the expected minimum response time for a set of $f + 1 - m$ replicas (also known as the first-order statistic) and $E[MAX_{f+1}]$ is the expected maximum response time of all $f + 1$ replicas, assuming all response times are identically distributed as some distribution Ψ.

Proof: During fault-free execution, the response time is given by the random variable MAX_{f+1}, i.e., the time to receive the latest of the $f + 1$ responses. Now, if up to $m \leq f$ replicas may be faulty, the response time of the earliest correct replica is given by the random variable MIN_{f+1-m}. The faulty replica(s) can delay the response time up to $K \cdot MIN_{f+1-m}$ without triggering a timeout. We must sum this fraction over all possible cases, $m = 1, 2, ...f$ faults, normalized by the probability of m faults, p^m. ∎

In the worst case, faulty replicas can inflate the average response time by up to a factor K, e.g., when $E[MAX_{f+1}]$ is not much greater than $E[\Psi]$ (because there is little variance in response times from different replicas) and there is only one correct replica, i.e., $m = f$. However, several factors mitigate this inflation in practice. First, if there is little variance in response times, then a small value of K suffices to keep the probability of false timeouts negligibly small. Second, in order to inflate response times without triggering a timeout, faulty replicas need to know the response time of the earliest response from a correct replica, which requires collusion with the agreement cluster. Third, the inflation is only for execution times and does not affect WAN propagation and transmission delays that typically account for the bulk of the client-perceived or end-to-end delays. Finally, as response times can be inflated only by faulty replicas, proactive recovery can further limit performance degradation under faults by limiting the length of time for which a faulty replica operates undetected [5].

9.2 Wakeup and Shutdown Rules

In this section we prove the following theorem about ZZ's wakeup and shutdown rules.

THEOREM 2 *If a wakeup occurs, ZZ will be able to terminate at least one faulty or slow replica.*

To ensure this theorem, ZZ uses the following wakeup rule:

Wakeup Rule: A wakeup happens if and only if a mismatch report is "blocking".

A mismatch occurs when an agreement replica receives execution replies which are not identical. Suppose that for a particular request there are $g + c$ agreement replicas which experience a mismatch. Consider the *mismatch matrix* of size (f+1)*(g+c) where entry i, j corresponds to the reply by execution replica E_i as reported by agreement node A_j. Let the execution mismatch of a row be defined as the smallest number of entries that need to be changed in order to make all $g + c$ entries in that row identical. Let the smallest such execution mismatch across all $f + 1$ rows be m. The mismatch is considered blocking if $m < c$.

To understand the difference, consider two examples where the client receives conflicting responses P and Q, and $f = g = 1$,

Full Matrix	**Mixmatch Matrix**	
$A_1 A_2 A_3 A_4$	$A_3 A_4$	
$E_1 : Q\ Q\ P\ Q$	$E_1 : P\ Q$	$c = 1$
$E_2 : Q\ Q\ Q\ P$	$E_2 : Q\ P$	$m = 1$

In this scenario, the mismatch matrix has size $g+c = 2$. Since $g = 1$, $c = 1$. Both rows require one entry to be changed in order to create a match, so the minimum mismatch is $m = 1$. Since $m = c$, this is not a blocking fault. The client will be able to receive a reply affirmation that Q is correct from A_1 and A_2. Note that if an additional node were woken up, it would not be possible to tell which execution node was faulty since it is impossible to tell if A_3 or A_4 is also faulty.

Full Matrix	**Mixmatch Matrix**	
$A_1 A_2 A_3 A_4$	$A_2 A_3$	
$E_1 : Q\ Q\ Q\ P$	$E_1 : Q\ Q$	$c = 1$
$E_2 : Q\ P\ P\ P$	$E_2 : P\ P$	$m = 0$

The second scenario illustrates a blocking mismatch. In this case, the *rows* in the mismatch matrix require no changes, thus $m = 0$. Since $m < c$ we have a blocking fault. This makes sense because there is no way to tell whether Q or P is the correct response, so a wakeup is required. To ensure Theorem 2, we also must guarantee:

Claim: A client will receive an affirmation certificate unless a mismatch is blocking.

Since $g+c$ agreement replicas report a mismatch, the number of replicas that have matching replies is $2g + 1 - c$. For a client to get an affirmation certificate, we must show that there are at least $g+1$ *correct* agreement replicas with matching replies whenever a blocking mismatch occurs. Since all rows have an execution mismatch of at least m, at least m of the agreement replicas out of the $g + c$ reporting a mismatch must be faulty. Thus, at most $g-m$ of the remaining $2g+1-c$ replicas can be faulty, so $2g+1-c-(g-m) = g+1-c+m$ are correct. If the fault is categorized as non-blocking, then $m >= c$, giving us at least $g + 1$ correct agreement replicas with matching replies. These nodes will be able to provide an execution affirmation to the client without requiring a wakeup, proving the claim.

Claim: A blocking mismatch implies that at least one execution replica other than a replica with the minimal execution mismatch is faulty.

We prove the claim by contradiction. Consider a replica with the minimal execution mismatch m and suppose that it is the only faulty execution replica. Since $g + c$ agreement replicas report a mismatch, at least $g + k - m$ agreement replicas must be faulty. However, if the mismatch is blocking, $m < k$, which implies that at least $g + 1$ agreement replicas must be faulty, resulting in a contradiction. Thus, there must be at least one other faulty execution replica.

Shutdown Rule: Shut down any f execution replicas not including a replica with the minimum execution mismatch.

Claim: The shutdown rule shuts down at least one faulty replica.

The proof follows from the claim above.

9.3 Replication Cost

ZZ provides the following guarantee about its expected replication cost:

THEOREM 3 *The expected replication cost of ZZ is less than* $(1 + r)f + 1$, *where* $r = 1 - (1 - p)^{f+1} + (1 - p)^{f+1}\Pi_1$.

where p is the probability of a node being faulty and Π_1 is the probability of a correct node causing a false timeout.

Lemma 1: The probability of a false timeout, Π_1 is bounded by:
$Pi_1 \leq \int_{t=0}^{\infty} \mathrm{p}[\Psi_{(1)} \leq t]\mathrm{P}[\Psi_{(f+1)} > Kt]dt$

Proof: A false timeout occurs when the timing assumptions are not satisfied, i.e., correct execution replicas return responses within times that vary by more than a factor K. The probability of a false timeout when the first response arrives within time t is given by $\mathrm{P}[\Psi_{(1)} \leq t] \cdot \mathrm{P}[\Psi_{(f+1)} > Kt \mid \Psi_{(1)} \leq t]$. The first term is the probability that the first response arrives within time t and the second term is the probability that the $(f+1)$th response arrives within time Kt conditional on the first response arriving within time t. Note that $\Psi_{(1)}$ and $\Psi_{(f+1)}$ are not independent, hence the conditional in the second term. The probability of a false timeout Π_1 is

$$\Pi_1 = \int_{t=0}^{\infty} p[\Psi_{(1)} \leq t] \cdot P[\Psi_{(f+1)} > Kt \mid \Psi_{(1)} = t]dt$$

where the integral is over all possible values of t and the notation $p[.]$ using a small 'p' denotes the pdf[2], i.e., $p[\Psi_{(1)} \leq$

[2]The notation $p_{\Psi_{(1)}}(t)$ is more common, but results in a confusing subscript overkill.

$t] = \frac{d}{dt}(P[\Psi_{(1)} \leq t])$. In order to prove the lemma, we note that removing the conditional in the second term inside the integral strictly increases the second term for all possible values of t. In other words, the probability that the $(f+1)$th response time is greater than Kt is greater than the probability that the $(f+1)$th response time is greater than Kt conditional on the first response time being equal to t. Thus,

$$\Pi_1 \leq \int_{t=0}^{\infty} p[\Psi_{(1)} \leq t] \cdot P[\Psi_{(f+1)} > Kt]dt$$

proving the lemma. ∎

Note that the above Lemma is only the probability of a false timeout at a single correct agreement replica. The probability of a new execution replica being woken up is even smaller as $g + 1$ agreement replicas must be able to produce a convictable timeout.

COROLLARY 1 *The probability of a new execution replica being unnecessarily woken up despite all execution replicas being correct is at most* Π_1.

The above lemma can be used to numerically compute the bound on the probability of a false timeout or wakeup. For example, if Ψ is exponential with mean $\frac{1}{\lambda}$, then $\Psi_{(1)}$ is also an exponential with mean $\frac{1}{(f+1)\lambda}$ with a pdf given by $p[\Psi_{(1)} \leq t] = \lambda(f+1)e^{-\lambda(f+1)t}$; and $\Psi_{(f+1)}$ is given by $P[\Psi_{(f+1)} > t] = 1 - (1 - e^{-\lambda t})^{f+1}$.

Given this bound on false timeouts and the probability of a node being faulty p, we can prove Theorem 3 as follows.

Proof: The replication cost is more than $f + 1$ only when a fault is detected either because some replica is faulty or because a false timeout occurs. The probability that at least one replica is faulty is $1 - (1 - p)^{f+1}$. When all replicas are correct, Lemma 1 implies that a false timeout occurs with probability at most Π_1. Thus, a fault is detected with probability $r \leq 1 - (1 - p)^{f+1} + (1 - p)^{f+1}\Pi_1$. Up to f replicas may be woken up when a fault is detected. Hence the theorem.