Velos: One-sided Paxos for RDMA applications

Rachid Guerraoui  
EPFL

Antoine Murat  
EPFL

Athanasios Xygkis  
EPFL

Abstract

Modern data centers are becoming increasingly equipped with RDMA-capable NICs. These devices enable distributed systems to rely on algorithms designed for shared memory. RDMA allows consensus to terminate within a few microseconds in failure-free scenarios, yet, RDMA-optimized algorithms still use expensive two-sided operations in case of failure. In this work, we present a new leader-based algorithm for consensus based on Paxos that relies solely on one-sided RDMA verbs. Our algorithm decides in a single one-sided RDMA operation in the common case, and changes leader also in a single one-sided RDMA operation in case of failure. We implement our algorithm in the form of an SMR system named Velos, and we evaluated our system against the state-of-the-art competitor Mu. Compared to Mu, our solution adds a small overhead of \( \approx 0.6 \mu s \) in failure-free executions and shines during failover periods during which it is 13 times faster in changing leader.

1 Introduction

RDMA is becoming increasingly popular in data centers. This networking technology is implemented in modern NICs and allows for Remote Direct Memory Access, where a server can access the memory of another without involving the CPU of the latter [24]. As a result, RDMA paves the way for distributed algorithms in the shared-memory model that are no longer confined within a single physical server.

Communication over RDMA takes two forms. One form is one-sided communication which shares similar semantics to local memory accesses: a server directly performs READs and WRITEs to the memory of a remote server without involving the remote CPU. The other form is called two-sided communication that has similar semantics to message passing: the server SENDs a message to the remote server which involves its CPU to process the RECEIVEd message.

A significant body of work on consensus [4] over RDMA has been conducted over the past decade [2, 14, 23, 25]. These solutions primarily focus on increasing the throughput and lowering the latency of common case executions, thus achieving consensus in the order of a few microseconds. Nevertheless, outside of their common case, these systems suffer from orders of magnitude higher failover times, ranging from 1ms [2] to tens or hundreds of ms [23, 25].

In this work we introduce a leader-based consensus algorithm based on the well-known Paxos [16] that is efficient both during the common-case as well as during failover. Our algorithm provides comparative performance to the state-of-the-art system Mu, i.e., it achieves consensus in under 1.9\( \mu s \), making it only twice as slow as Mu. At the same time, our algorithm is 13 times faster than Mu in the event of a leader failure and manages to failover in under 65\( \mu s \). To achieve this, our algorithm relies on one-sided RDMA WRITEs as well as Compare & Swap (CAS), a capability that is present in RDMA NICs.

The basic idea behind our algorithm is that the original Paxos algorithm contains RPCs that are simple enough to be replaced with CAS operations. The CAS operations are initiated by the leader and are executed on the memory of the participating consensus nodes (leader and followers). We first describe a version of our algorithm for single-shot consensus, where we provide proofs of correctness. Then, we continue by extending this version to a fully-fledged system, in which our algorithm takes a form similar to multishot Paxos [17]. In the case of a stable leader, it decides in a single CAS operation to a majority. In the event of a failure, the leader is changed in a single additional CAS operation to a majority.

The rest of this document is as follows: In section 2, we present related work. In section 3, we introduce the consensus problem and present the well-known Paxos algorithm. In section 4, we explain how to transform the RPC algorithm into a CAS-based one and we prove the correctness of this transformation. In section 5, we discuss further practical considerations, which are relevant in converting our single-shot consensus algorithm to a multi-shot one. In section 6, we discuss our implementation. In section 7, we evaluate the performance of our solution against Mu, the most recent
state-of-the-art system that implements consensus.

2 Related work

Mu

Mu [2] is an RDMA-powered State Machine Replication [3] system that operates at the microsecond scale. Similarly to our system, it decides in a single amortized RDMA RTT. This is achieved by relying on RDMA permissions [24]. Mu ensures that at any time, at most one process can write to a majority and decide, which ensures consensus’ safety. In case of leader failure, Mu requires permission changes that take \(\approx 250\mu s\). Mu thus fails at guaranteeing microsecond decisions in case of failure.

APUS

APUS [25] is a Paxos-based SMR system. It was tailored for RDMA. It doesn’t use expensive permission changes but relies heavily on two-sided communication schemes. While it provides short failovers, its consensus engine involves heavy CPU usage at replicas and is significantly slower than Mu’s.

Disk Paxos

Disk Paxos [10] observes that Paxos’ acceptors can be replaced by moderately smart and failure-prone shared memories. Their work can be done by proposers as long as they are able to run atomic operations at the shared resource. This work is purely theoretical.

3 Preliminaries

In this section, we state the process and communication model that we assume for the rest of this work. Then we formally introduce the problem of consensus and we present the well-known Paxos algorithm.

3.1 Assumptions

We consider the message-and-memory (M&M) model [1], which allows processes to use both message-passing and shared-memory. Communication is assumed to be lossless and provides FIFO semantics. The system has \(n\) processes \(\Pi = \{p_1, \ldots, p_n\}\) that can attain the roles of proposer or acceptor. In the system, we assume that there are \(p\) proposers and \(n\) acceptors, where \(1 < p < |\Pi|\). Processes can fail by crashing. Up to \(p - 1\) proposers and \(\left\lceil \frac{n-1}{2} \right\rceil\) acceptors may fail. As long as a process is alive, its memory is remotely accessible. When a process crashes, its memory also crashes. In this case, subsequent memory operations do not return a response. The system is asynchronous in that it can experience arbitrary delays.

3.2 Consensus

In the consensus problem, processes propose individual values and eventually irrevocably decide on one of them. Formally, Consensus has the following properties:

Uniform agreement If processes \(i\) and \(j\) decide \(val\) and \(val’\), respectively, then \(val = val’\).

Validity If some process decides \(val\), then \(val\) is the input of some process.

Integrity No process decides twice.

Termination Every correct process that proposes eventually decides.

It is well known that consensus is impossible in the asynchronous model [9]. To circumvent this impossibility, an additional synchrony assumption has to be made. Our consensus algorithm provides safety in the asynchronous model and requires partial synchrony for liveness. For pedagogical reasons and in order to facilitate understanding, we implement our consensus algorithm by merging together the following abstractions:

- Abortable Consensus [4], an abstraction weaker than Consensus that is solvable in the asynchronous model,

- Eventually Perfect Leader Election [6], which relies on the weakest failure detector required to solve Consensus.

3.3 Abortable Consensus

Abortable consensus has the following properties:

Uniform agreement If processes \(i\) and \(j\) decide \(val\) and \(val’\), respectively, then \(val = val’\).

Validity If some process decides \(val\), then \(val\) is the input of some process.

Termination Every correct process that proposes eventually decides or abort.

Decision If a single process proposes infinitely many time, it eventually decides.

Algorithm 1 solves Abortable Consensus and is based on Paxos. Processes are divided into two groups: proposers or acceptors. Proposers propose a value for decision and acceptors accept some proposed values. Once a value has been accepted by a majority of acceptors, it is decided by its proposer.

Intuitively, the algorithm is split in two phases: the Propose phase and the Accept phase. During these phases, messages from the proposer are identified by a unique proposal number. The Prepare phase serves two purposes. First, the
Algorithm 1: Abortable Consensus

Proposers execute:
upon (Init):
    decided = False
    proposal = id
    proposed_value = ⊥
propose(value):
    proposed_value = value
    if not decided:
        if prepare():
            accept()
prepare():
    proposal = proposal + |I|
    broadcast (Prepare | proposal)
    wait for a majority of (Prepared | ack, ap, av)
    if any av returned, replace proposed_value with av with ⩾ highest ap
    if any not ack:
        trigger (Abort)
    return False
accept():
    broadcast (Accept | proposal, proposed_value)
    wait for a majority of (Accepted | mp)
    if any mp > proposal:
        trigger (Abort)
    if leader and not proposed:
        while True:
            proposed_value = value
    else:
        decided = true
        trigger (Decide | proposed_value)

Acceptors execute:
upon (Init):
    min_proposal = 0
    accepted_value = ⊥
upon (Prepare | proposal):
    min_proposal = n
    reply (Prepared | min_proposal == n, accepted_proposal, ⩾ accepted_value)
upon (Accept | proposal, value):
    if proposal ≥ min_proposal:
        accepted_proposal = min_proposal = n
        accepted_value = value
        reply (Accepted | min_proposal)

3.4 From Abortable Consensus to Consensus

One can solve Consensus by combining Abortable Consensus together with Eventually Perfect Leader Election (Ω). In Abortable Consensus a proposer is guaranteed to decide, rather than abort, if it executes unobstructed. The role of Ω is to ensure this condition is ensured. Informally, Ω guarantees that eventually all correct proposers will consider a single one of them to be the leader. As long as a proposer is considering itself as the leader it keeps on proposing its value using Abortable Consensus. Eventually, Ω will mark a single correct proposer as the leader, which will try to propose unobstructed and decide. The leader can then broadcast the decision to the rest of the proposers. Algorithm 2 provides the implementation of this idea. A proof of its correctness is given in [4].

Algorithm 2: Consensus from Abortable Consensus

Proposers execute:
upon (Init):
    proposed_value = ⊥
    leader = proposed = decided = False
upon (Trust | p):
    if p. ⩾ self then leader = True
    else leader = False
propose(value):
    proposed_value = value
    while True:
        if leader and not proposed:
            trigger (AbortableConsensus, Propose | ⩾ proposed_value)
    upon (AbortableConsensus, Decide | value):
        broadcast (Decided | value)
    upon (AbortableConsensus, Abort):
        proposed = False
    upon (Decided | value):
        if not decided:
            decided = True
            trigger (Decide | value)

Notice that Abortable Consensus differs from Consensus only in its liveness property. The former is essentially an obstruction free consensus implementation in which no proposer may decide under contention. However, Abortable Consensus retains the safety properties of Consensus. Thus, for the rest of this work we will concentrate on Abortable Consensus and we will transform it into a CAS-based algorithm.

4 One-sided Consensus

In this section, we explain how to transform the two-sided algorithm presented in section 3.3 into a one-sided one. To do so, we first establish the equivalence between the RPCs used in algorithm 1 and CAS. Then, we take advantage of this equivalence and replace RPCs with CAS in algorithm.
1. The resulting CAS-based Abortable Consensus is given in algorithm 4.

### 4.1 One-sided obstruction-free RPC

Observe that algorithm 1 uses message passing (i.e., RPC) in a very specific form. The acceptors keep track of only three variables: min\_porposal, accepted\_proposal and accepted\_value. In both the Accept and the Prepare phases, acceptors atomically update these values upon a very simple condition (i.e., a simple comparison) and return some of them. In this section, we propose and prove a simple obstruction-free transformation to turn such RPCs into purely one-sided conditional writes using CAS.

```
Algorithm 3: CAS-based RPC

rpc(x):
  if compare(x, state):
    state = f(state, x)
    return projection(state)
  return projection(f(state, x))
```

In algorithm 3, we assume that the whole state of a process (i.e., all its variables) is stored in state. In the case of RPC (line 1), the caller sends x to the callee. The callee deterministically compares x with its state using compare. If the comparison succeeds, its state is deterministically mutated using the function f. In any case, the callee extracts part of its state using projection and returns it to the caller. By convention, rpc runs atomically.

We prove below that if the callee’s state is accessible by the caller via shared memory, and compare, f, projection are known to the caller, then rpc and cas-rpc are strictly equivalent except in the case where cas-rpc aborts.

**Lemma 4.1.** If cas-rpc does not abort, rpc and cas-rpc are equivalent.

**Proof.** An execution of rpc solely depends on the value of state and the input value x. We denote such execution of rpc with \langle state, x\rangle_{rpc}. If an execution of cas-rpc does not abort, it solely depends on the value of expected fetched at line 7 and the input value x. We denote such execution of cas-rpc with \langle expected, x\rangle_{cas-rpc}.

We show that any execution \langle s, x\rangle_{rpc} is equivalent to the execution \langle s, x\rangle_{cas-rpc} in the sense that both rpc and cas-rpc will have the same value of state and return the same projection at the end of their execution.

If an execution \langle s_1, x\rangle_{rpc} makes the comparison at line 2 fail, then state is not modified and projection(s_1) is returned. In the execution \langle s_1, x\rangle_{cas-rpc}, the comparison at line 8 will also fail (as the comparison is deterministic) and projection(s_1) is also returned without modifying the remote state. In this case, both executions are equivalent.

If an execution \langle s_2, x\rangle_{rpc} makes the comparison at line 2 succeed, then state is modified to f(s_2, x) and projection(f(s_2, x)) is returned. In the execution \langle s_2, x\rangle_{cas-rpc}, the comparison at line 8 will also succeed (since the comparison is deterministic). As the run is assumed to not abort, the CAS will succeed. Thus the remote state will atomically be updated from s_2 to f(s_2, x) and f(s_2, x) is also returned. So, in this case, both executions are also equivalent.

In addition, this transformation is safe in case of obstruction.

**Lemma 4.2.** If cas-rpc aborts, it has no side effect.

**Proof.** If cas-rpc aborts, the comparison at line 13 failed. This in turn implies that the CAS at line failed and thus that state is unaffected by the execution.

From lemmas 4.1 and 4.2, cas-rpc exhibits all-or-nothing atomicity. We now prove that such a transformation in obstruction-free.

**Lemma 4.3.** If cas-rpc runs alone (i.e., unobstructed), it does not abort.

**Proof.** Let’s assume by contradiction that cas-rpc runs alone and aborts. For cas-rpc to abort, the comparison at line 13 must fail. This in turn implies that the CAS at line 12 failed, i.e., that the current value of state does not match the expected one. For this to happen the state must have been updated between line 7 and 12 by another process. This means that there was a concurrent execution, a contradiction.

### 4.2 A purely one-sided consensus algorithm

Algorithm 4 implements Abortable Consensus by replacing RPCs in algorithm 1 with the one-sided obstruction-free RPCs introduced in 4.1. As this new algorithm relies on CAS for comparisons, it aborts in situations where the original algorithm would have succeeded. For example, consider the following execution: Let proposers P_1 and P_2 concurrently initiate the Prepare phase with respective proposals 1 and 2. Both fetch the remote state and get \langle 0, 0, \bot\rangle. Then, P_1 succeeds at writing its proposal to acceptor A_1. Later on, the CAS of P_2 fails at A_1 as the value is now \langle 1, 0, \bot\rangle instead of the expected \langle 0, 0, \bot\rangle. Thus, P_2 aborts even if it had a larger proposal number than P_1. The more relaxed comparison in the original algorithm would not have caused P_2 to abort.
Algorithm 4: CAS-based Abortable Consensus

Proposers execute:
upon (Init):
  decided = False
  proposal = id
  proposed_value = ⊥
propose(value):
  proposed_value = value
  if not decided:
    if prepare():
      accept()
    return
  prepare():
    proposal = proposal + |I|
  execute in parallel cas_prepare(p, proposal) for p in Acceptors
  wait for a majority to abort or return (ack, ap, av)
  if any returned, replace proposed_value with av with
  ⊥ → highest ap
  if any aborted or not ack:
    trigger (Abort)
    return False
  return True
accept():
  execute in parallel cas_accept(p, proposal, value)
  for p in Acceptors
  wait for a majority to abort or return mp
  if any aborted or returned mp > proposal:
    trigger (Abort)
    return
  else:
    decided = True
    trigger (Decide | proposed_value)
cas_prepare(p, proposal):
  expected = fetch_state(p)
  if not proposal > expected.min_proposal:
    return (false, expected.accepted_proposal, expected.
    ⊥ → accepted_value)
  move_to = expected
  move_to.min_proposal = proposal
  old = cas(p.state, expected, move_to)
  if old == expected:
    return (true, expected.accepted_proposal, expected.
    ⊥ → accepted_value)
  abort()
cas_accept(p, proposal, value):
  expected = fetch_state(p)
  if not proposal > expected.min_proposal:
    return expected.min_proposal
  move_to = expected
  move_to.min_proposal = proposal
  move_to.accepted_proposal = proposal
  move_to.accepted_value = value
  old = cas(p.state, expected, move_to)
  if old == expected:
    return expected.min_proposal
  abort()

Acceptors execute:
upon (Init):
  state = {min_proposal, accepted_proposal, accepted_value ⊥}

Lemma 4.4. Algorithm 4 preserves the decision property of Abortable Consensus.

Proof. If a single process proposes infinitely many time, it will eventually run one-sided RPCs obstruction-free. By lemma 4.3, this guarantees that eventually the one-sided RPCs will terminate without aborting. In such case, lemma 4.1 guarantees the execution to be equivalent to one of the original algorithm. Thus, the transformation preserves the decision property of the original algorithm.

Lemma 4.5. Algorithm 4 preserves the termination property of Abortable Consensus.

Proof. Assuming a majority of correct acceptors, all CASes will eventually return or abort. Due to the absence of loops or blocking operations (apart from waiting for a reply from a majority of acceptors), a proposer that invokes propose will either abort or decide.

The only execution difference between both algorithms is that some executions of the transformed algorithm may abort, where the original one would not. Nevertheless, aborting does not violate safety.

Lemma 4.6. Algorithm 4 preserves the safety properties of Abortable Consensus.

Proof. Assume by contradiction that adding superfluous aborts in the original algorithm can violate safety. In a first execution $E_1$, processes $\{P_1, ..., P_n\}$ deviate from the original algorithm and abort (or decide) at times $\{t_1, ..., t_n\}$ after which the global state is $\{S_1, ..., S_n\}$. At some point, safety is violated. In a second execution $E_2$, processes $\{P_1, ..., P_n\}$ crash at times $\{t_1, ..., t_n\}$ after which the global state is $\{S_1, ..., S_n\}$. As the original algorithm tolerates arbitrarily many proposer crash failures, safety is not violated. Proposers cannot distinguish both executions. Thus, safety cannot be violated, hence a contradiction. Thus, adding superfluous aborts preserves safety and algorithm 4 preserves safety.

Theorem 4.7. By lemmas 4.5, 4.5, 4.6, algorithm 4 implements Abortable Consensus.

4.3 Streamlined one-sided algorithm

Section 4.2 introduced a simple one-sided consensus algorithm built by replacing RPCs with weaker (i.e., abortable) one-sided RPCs and proved its correctness. The resulting algorithm exhibits flagrant inefficiencies that can be fixed to reduce the number of network operations to 2 CASes in the common case.

First, it is not required to fetch the remote state at the start of each RPC. As it is safe to have stale expected states, it is safe to use optimistic predicted states deduced from previous CASes. Predicted states can thus be initialized to $\{0, 0, \perp\}$ and updated each time a CAS complete (either succeeding or not). Moreover, wrongly predicting states can only result in superfluous aborts which have been proven to be safe by lemma 4.6. Thus, it is safe to optimistically assume that on-flight CAS will succeed.
Second, in the Prepare phase, proposal can be increased upfront to be higher than any predicted remote min_proposal to reduce predictable abortions.

Algorithm 5 gives the algorithm obtained after applying these optimisations.

Algorithm 5: Streamlined One-sided Abortable Consensus

Proposers execute:

upon (Init):
  predicted = {min_proposal: 0, accepted_proposal: 0, accepted_value: ⊥}
  decided = False
  proposal = id
  proposed_value = ⊥
propose(value):
  proposed_value = value
  if not decided:
    if prepare():
      accept()
predicted:
  if reads[p] ∈ {predicted[p], ⊥}:
    predicted[p] = moves[p]
    #(lines 26-27)
  else:
    predicted[p] = reads[p]
  if any CAS failed (predicted[p] ≠ reads[p]):
    trigger (Abort)
    return false
  proposed_value = predicted[.] accepted_value with highest
  → accepted_proposal or proposed_value if none
  return true
accept():
  move_to = {proposal, proposal, proposed_value}
  for p in Acceptors:
    async reads[p] = cas(slot_p, predicted[p], move_to[p])
    wait until majority of slots are read
  if any CAS failed:
    for p in Acceptors:
      if reads[p] ∈ {predicted[p], ⊥}:
        predicted[p] = move_to[p]
      else:
        predicted[p] = reads[p]
    trigger (Abort)
    return
  decided = true
  trigger (Decide | proposed_value)

Acceptors execute:

upon (Init):
  state = {min_proposal: 0, accepted_proposal: 0, accepted_value: ⊥}

Said optimisations also preserve liveness. Let’s assume that a single proposer runs infinitely many time. Eventually, it will run obstruction-free. In the worst case, it will each time abort at line 34 or 52 because of a single wrongly guessed remote state and update its prediction. The optimistic update of expected states at lines 29 and 49 and the FIFO property of links provide that, once a remote state is correctly guessed, any later CAS will succeed. Thus, after at most \( n \) runs all CASes will succeed and the proposer will decide.

5 Practical Considerations

So far we have showed a new algorithm for doing RDMA-based consensus using CAS. Our algorithm presents a single instance of consensus, however most practical systems require to run consensus over and over.

State Machine Replication (SMR) replicates a service (e.g., a key-value storage system) across multiple physical servers called replicas, such that the system remains available and consistent even if some servers fail. SMR provides strong consistency in the form of linearizability [11]. A common way to implement SMR, which we adopt in this paper, is as follows: each replica has a copy of the service software and a log. The log stores client requests. We consider non-durable SMR systems [13, 15, 18, 19, 21, 22], which keep state in memory only, without logging updates to stable storage. Such systems make an item of data reliable by keeping copies of it in the memory of several nodes. Thus, the data remains recoverable as long as there are fewer simultaneous node failures than data copies [23].

A consensus protocol ensures that all replicas agree on what request is stored in each slot of the log. Replicas then apply the requests in the log (i.e., execute the corresponding operations), in log order. Assuming that the service is deterministic, this ensures all replicas remain in sync. We adopt a leader-based approach, in which a dynamically elected replica called the leader communicates with the clients and sends back responses after requests reach a majority of replicas. As already stated, we assume a crash-failure model: servers may fail by crashing, after which they stop executing.

Velos is the implementation of Algorithm 5 in the form of SMR. Implementing Velos leads to practical considerations and additional hardware-specific optimizations that are not present in Algorithm 5.

5.1 Pre-preparation

Unfortunately, our CAS approach prevent us from using the multi-Paxos optimisation [5, 17, 20]. Thus, each consensus slot must be prepared individually. Nevertheless, a leader can prepare slots in advance so that the it can decide running only the accept phase on its critical path. In this case, the leader decides in a single CAS RTT. Doing so requires a stable leader. Switching to another leader requires re-preparing slots. Fortunately, this will most likely succeed in a single
CAS RTT by optimistically predicting remote slots to have been prepared by the failed leader.

5.2 Limited CAS size

Algorithm 5 assumes that a state can fit within a single CAS. Current RDMA NICs support CAS up to 8 bytes. As both min_proposal and accepted_proposal must be the same size, both fields are limited to at most 31 bits with 2 bits being dedicated to storing the accepted_value.

One may be afraid of proposal fields overflowing (either breaking safety or decision). In such an extremely unlikely case (and actually more for completeness), the abstraction can fallback to traditional RPC. Switching to RPC likely case (and actually more for completeness), the abstraction can fallback to traditional RPC. Switching to RPC can be safely implemented as follows: Once the RDMA-exposed min_proposal of an acceptor reaches $2^{31} - |\Pi|$, every proposer switches to RPC to communicate with this specific acceptor. At a regular time interval, acceptors check their state and, if above the threshold, execute algorithm 1 with min_proposal, accepted_proposal and accepted_value variables initiated to match state.

Another concern is the limited accepted_value field size. Depending on the upper application, the value can be limited within the CAS. Otherwise, a simple indirection can be put in place. Instead of deciding on the proposed value itself, one can decide on the proposer’s id. This can be achieved by RDMA-writing the actual value to a majority of acceptors in a dedicated write-exclusive slot before running the accept phase. RDMA RC QP FIFO semantics can be leverages to do so at minimum cost. The write WQE can be prepended to the Accept WQE, made unsignaled, and posted with Doorbell batching. If the CAS completes, then the value was written at a majority of acceptors and will always be recoverable.

5.3 Device Memory

Modern RNICs allow RDMA exposure of their own internal memory. This feature is called Device Memory (DM). In Mellanox’ Connect-X6, the available memory is slightly larger than 100KiB. RDMAing this memory is faster than accessing the main memory as it removes an extra PCIe communication from the critical path. All RDMA verbs benefit from a speedup. Moreover, DM reduces atomic verbs contention. As one-sided Paxos makes acceptors fundamentally passive, DM can be used without additional cost. DM can also be leveraged when RDMA-writing the actual value as described in the previous section to PCIe-transfer the payload only once.

5.4 Piggybacking decisions

Consensus slots can be augmented with an additional previous_decision value. This way, if every node endorses both the proposer and the acceptor roles, it can learn that a value has been decided for slot $s - 1$ simply by reading its local slot $s$. With this strategy, values are decided in a single CAS and decisions learned with no additional communication in the common case.

5.5 Unavailable features

Modern RNICs lack some features and performances that would make one-sided Paxos even more appealing. We believe that these features can reasonably be expected to be provided by future RNICs.

First are global CASes. CX6 atomics are only guaranteed to be atomic with other operations executed by the same HCA. This prevent us from using CPU-side CASes to update the state of the proposer, which could save one MMIO ($\approx 300\text{ns}$).

Second are unaligned CASes. Currently, CASes cannot overlap 2 consecutive slots. Such a feature would allow a proposer to run the Accept phase for slot $s$ and the Prepare phase for slot $s + 1$ in a single CAS, definitely removing the need for batch pre-preparation from the critical path.

Third, even using Device Memory, CAS exhibits a latency twice higher than RDMA WRITEs, which makes Mu twice as fast in the failure-free scenario.

6 Implementation

Our algorithm is implemented in 1046 lines of C++17 code (CLOC [7]). It uses the ibverbs library for RDMA over Infiniband and it relies on the reusable abstractions provided by the source code of Mu. Furthermore, we implemented the optimizations mentioned in section 5, apart from “Piggybacking decisions”.

We implement a new leader election algorithm that departs from Mu’s design. Our algorithm relies on modifications of the Linux kernel and is composed of two separate modules, namely the interceptor and the broadcaster module. In Linux, when a process crashes control is transferred to the kernel which takes care of cleaning up the process. Our kernel modifications allow processes to instruct the kernel—via a prctl system call—that upon failure the interceptor module should be invoked during the cleaning up. The interceptor module is then responsible for invoking the broadcaster module. The broadcasting module broadcasts a message saying that the process has crashed. This message is picked up by correct processes that subsequently stop trying to communicate with the crashed process. Our kernel modifications and kernel modules implementation span in 352 lines of C code.
7 Evaluation

Our goal is to evaluate whether our implementation can achieve consensus within a few microseconds and whether it handles failures with the least amount of delay. Concretely, with our evaluation we aim at answering the following questions:

- Does our implementation imposes a small overhead compared to Mu in the common case?
- What is the total failover time during a leader crash?

We evaluate our system on a 3-node cluster, the details of which are given in Table 1.

In the reported numbers we show 3-way replication, which accounts for most real deployments [12]. In all of our experiments we ignore the existence of a client issuing requests to our system.

We measure time using the POSIX `clock_gettime` function, with the `CLOCK_MONOTONIC` parameter. In our deployment, the resolution and overhead of `clock_gettime` is around 16–20 ns [8].

Table 1: Hardware details of machines.

| Parameter          | Details                                      |
|--------------------|----------------------------------------------|
| CPU                | 2x Intel(R) Xeon(R) Gold 6244 CPU @ 3.60GHz  |
| Memory             | 2x 128GiB                                    |
| NIC                | Mellanox ConnectX-6                          |
| Switch             | Mellanox MSB7700 EDR 100 Gbps               |
| Links              | 100 Gbps Infiniband                          |
| OS                 | Ubuntu 20.04.2 LTS                           |
| Kernel             | 5.4.0-74-custom                              |
| RDMA Driver        | Mellanox OFED 5.3-1.0.0.1                    |

We measure time using the POSIX `clock_gettime` function, with the `CLOCK_MONOTONIC` parameter. In our deployment, the resolution and overhead of `clock_gettime` is around 16–20ns [8].

7.1 Common-case Replication Latency

We first evaluate the latency of our system under no leader failure. We measure the latency at the leader. Latency refers to the time it takes for `propose` of Algorithm 5 to execute. We show the time it takes to replicate messages of different sizes in Figure 1.

For messages of 1 byte, Velos replicates a request using only a RDMA CAS. On the other side, Mu always uses a single RDMA WRITE to replicate. For messages up to 128 bytes Mu manages to inline the message in the RDMA request, thus exhibiting almost constant replication latency. RDMA WRITE and RDMA CAS are both one-sided operations, but they have different latency. Sending 3 RDMA CASes and waiting for a majority of replies costs around 1.9µs, while the same communication pattern with RDMA WRITEs costs 1.25µs. This is seen from the time difference for 1B messages.

For larger payload sizes, Velos replicates a request using a RDMA CAS and an additional RDMA WRITE. In other words, Velos does exactly what Mu does, apart from having an extra RDMA CAS for every replicated message. Given that the latency of RDMA CAS is constant (in the absence of CAS-contention, i.e., when having a stable leader), the impact of the overhead in the replication latency of Velos compared to Mu diminishes for larger message sizes.

Figure 1 also demonstrates the effect of using Device Memory. When Velos relies exclusively on Device memory for its RDMA WRITEs and RDMA CASes, it gains 200ns on replication latency.

7.2 Fail-over Time

During a leader crash, Velos replicas receive the failure detection event and a new leader is elected. The new leader immediately starts replicating new messages among itself and the remaining replica \( R \).

Figure 2 evaluates the time it takes for \( R \) to discover a new replicated value in its log. When a stable leader replicates requests, the throughput is at around 42 requests per 100µs. In other words, \( R \) discovers a new entry in its log approximately every 2.5µs. When the leader fails the throughput drops to 0 and replica \( R \) discovers a new value after approximately 65µs. The subsequent few replication requests from the new leader take between 3µs to 3.6µs, which is evident in the throughput curve. When the new leader takes over, the throughput curve exhibits a not-so-steep trajectory, before stabilizing again at the throughput of 42 requests per 100µs. This is because the new leader’s cache is cold and initial replication requests result in higher replication latency. Soon after the new leader manages to replicate new requests in approximately 2.5µs.

Comparing Velos to Mu, the latter faster than Velos in the common case but slower during leader change. Mu manages to replicate requests using a single WRITE but it relies on permissions to handle concurrent leaders during leader fail-
The cost of changing permissions, as presented in Mu, is approximately 250 $\mu s$ just for changing the leader. Mu requires an additional 600 $\mu s$ to detect the leader failure. On the other side, Velos requires approximately 30 $\mu s$ to detect a leader failure and an additional 35 $\mu s$ for the new leader to successfully replicate the next request. In other words, Velos is approximately 1.2 or 1.5 $\mu s$ slower than Mu in the common case, depending on whether it relies or not in Device Memory. Velos is significantly faster than Mu during leader change, as it detects a leader failure 20 times faster than Mu and replicates the next request 7 times faster than Mu. Overall, Velos is 13 times faster than Mu during leader change.

**8 Conclusion**

Consensus is a classical distributed systems abstraction that is widely used in the data centers. RDMA enables consensus to achieve lower decision latency not only due to its intrinsic latency characteristics as a network fabric, but also due to its semantics. RDMA semantics such as permission changes improve the latency of consensus in the common case, as they enable consensus to decide by using a single RDMA WRITE. However, the non-common case during which failures still exhibit latency in the order of a millisecond.

Velos is a state machine replication system that can replicates requests in a few microseconds. It relies a modified Paxos algorithm that replaces RPCs (i.e. message passing) with RDMA Compare-and-Swap. As a result, Velos exhibits common-case latency that is competitive to Mu’s latency and shines during failure. Velos manages to switch to a new leader and start replicating new requests in under 65 $\mu s$, meaning that Velos has an extra 9 of availability for the same number of expected failures, compared to Mu.

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