FireLedger: A High Throughput Blockchain Consensus Protocol *

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Abstract

Blockchains are distributed secure ledgers to which transactions are issued continuously and each block
of transactions is tightly coupled to its predecessors. Permissioned blockchains place special emphasis
on transactions throughput. In this paper we present FireLedger, which leverages the iterative nature
of blockchains in order to improve their throughput in optimistic execution scenarios. FireLedger trades
latency for throughput in the sense that in FireLedger the last \( f + 1 \) blocks of each node’s blockchain
are considered tentative, i.e., they may be rescinded in case one of the last \( f + 1 \) blocks proposers was
Byzantine. Yet, when optimistic assumptions are met, a new block is decided in each communication
step, which consists of a proposer that sends only its proposal and all other participants are sending a
single bit each. Our performance study demonstrates that in a single Amazon data-center, FireLedger
running on 10 mid-range Amazon nodes obtains a throughput of up to 160K transactions per second
for (typical Bitcoin size) 512 bytes transactions. In a 10 nodes Amazon geo-distributed setting with
512 bytes transactions, FireLedger obtains a throughput of 30K tps. Moreover, on higher end Amazon
machines, FireLedger obtains 20%−600% better throughput than state of the art protocols like HotStuff
and BFT-SMaRt, depending on the exact configuration.

1 Introduction

A blockchain is a distributed secure replicated ledger service designed for environments in which not all nodes
can be trusted [63]. Specifically, a blockchain maintains a distributed ordered list of blocks (the “chain”) in
which every block contains a sequence of transactions as well as authentication data about the previous
blocks; the latter typically relies on cryptographic methods. Hence, any attempt to modify part of the
blockchain can be detected, which helps to ensure the stability and finality of blockchain prefixes. Notice
that transactions may in fact be any deterministic computational step, including the execution of smart
contracts code. A primary challenge in implementing a blockchain abstraction is deciding on the order of
transactions in the chain.

Largely speaking, blockchains can be characterized as either unpermissioned or permissioned. In unper-
missioned blockchains, any node is allowed to participate in the computational task of deciding the ordering
of transactions and blocks as well as in the task of maintaining the blockchain’s state [44, 63]. In such
an environment, there is no trust between nodes and in fact, with crypto-currencies, participants have an
a-priori incentive to cheat. This implies utilizing significant cryptographic mechanisms to compensate for
this zero trust model, which limits the system’s throughput.

In contrast, in permissioned mode, the blockchain is executed among a set of \( n \) known participants
under the assumption that at most \( f \) of them are faulty [28]. For example, consider a consortium of
insurance companies each donating a node in order to maintain a common blockchain of insurance policies
and insurance claims. In this setting, blockchain becomes a special case of traditional replication state
machine (RSM) [55, 69]. A common approach to implementing RSM is by repeatedly running a consensus

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protocol to decide on the next transaction to be executed \[56\] with the optimization of batching multiple transactions in each invocation of the consensus protocol \[47\].

The assumed possible type of failures affects the type of consensus protocols that are used. **Benign** failures such as a node crash and occasional message omission can be overcome by benign consensus protocols, e.g., \[31, 51, 56, 64\]. On the other hand, **Byzantine** failures \[57\] in which a faulty node may arbitrarily deviate from its code require Byzantine fault tolerant protocols (BFT), e.g., \[19, 29, 53\]. The type of failures along with the synchrony assumptions about the environment in which the blockchain is implemented impact the minimal ratio between \(f\) and \(n\) required to enable a solution to the problem. For example, assuming a synchronous environment, benign consensus requires \(f < n\) while Byzantine consensus requires \(f < n/2\) in the signed messages model, and \(f < n/3\) in the oral messages model \[57\]. In the totally asynchronous case, the seminal FLP result showed that even benign consensus cannot be solved \[45\]. Yet, when enriching the environment with some minimal eventual synchrony assumptions, e.g., partial synchrony \[38, 41\], or with unreliable failure detector oracles \[32\], benign consensus becomes solvable when \(f < n/2\) and Byzantine consensus requires \(f < n/3\). This is as long as the network does not become partitioned \[21\].

Here we focus on permissioned blockchains assuming Byzantine failures and partial synchrony. According to a recent survey by PwC \[12\], only a third of companies currently using or planning to use blockchains intend on using unpermissioned blockchains. Also, many recent unpermissioned proposals, e.g., Algorand \[49\], Tendermint \[25\], HoneyBudger \[60\], and Thunderella \[65\], can be viewed as running a permissioned protocol coupled with a higher level meta-protocol that continuously selects which nodes can participate in the internal permissioned one. Variants of this approach are sometimes referred to as delegated proof of stake (dPoS) and **Proof of Authority** (PoA). Hence, any improvement in permissioned protocols will likely yield better unpermissioned protocols as well. As mentioned above, most solutions in this domain involve repeatedly invoking a consensus instance in order to decide on the next transaction or batch of transactions \[31, 56\]. This is vulnerable to performance attacks, as been identified in \[15\], many of which can be ameliorated by rotating the role of the consensus initial proposer on each invocation of the protocol \[15, 33, 72\].

Many of the above works try to optimize performance in the “common case” in which there are no failures and the network behaves in a synchronous manner. These situations are likely to be common in permissioned blockchains, e.g., executed between major financial institutions, established business partners, etc. Yet, all the above mentioned works run each protocol instance for completion. Alas, we claim that in a production blockchain system, where transactions are being submitted continuously for as long as the service exists, there is potential for reducing the per transaction and per block communication overhead. This is by assuming optimistically that the initial proposer of each consensus invocation is correct, and only performing a recovery phase periodically for a batch of affected consensus invocations and only if it is needed.

### Our Contributions

We propose FireLedger, a new communication frugal optimistic permissioned blockchain protocol. FireLedger utilizes the rotating proposer scheme while optimistically assuming that the proposer is correct and that the environment behaves synchronously. If these assumptions are violated, we do not insist on enforcing agreement immediately. Instead, we rely on the fact that at least one out of every \(f + 1\) proposers is correct. When a correct node discovers, using blockchain’s authentication data, that any of the last \(f + 1\) blocks was not decided correctly in the initial transmission phase, it runs a combined recovery phase for all these incorrectly executed invocations. This is by invoking a full Byzantine consensus protocol. At the end of this combined recovery phase, it is ensured that the current prefix of the blockchain is agreed by all correct nodes and will never change as long as there are at most \(f\) Byzantine failures. A single recovery phase may decide the last \(f\) blocks, thereby amortizing its cost.

The main benefit of our approach is that when the optimistic assumption holds, the communication overhead of deciding on a block involves a single proposer broadcasting its block and all other nodes broadcasting a single bit of unsigned protocol data. Further, a new block is being decided in each communication step. This is by leveraging the iterative nature of blockchain as well as the authentication data that is associated with each block header.

Notice that \(f\) is an upper bound on the maximal number of Byzantine nodes in the system. Yet, in
many permissioned blockchains settings, nodes are likely to be highly secured. Further, in our protocol any Byzantine deviation from the protocol results in a strong proof of which node was the culprit. Hence, we expect that in real deployments the optimistic assumptions will hold almost always. In particular, once a proof of Byzantine behavior is being generated, the corresponding Byzantine node will be removed from the system, often resulting in financial penalties and loss of face for the owner of this node.

The price paid by our algorithm is that finality of a decision is postponed for $f+1$ invocations (or blocks). That is, we trade bandwidth and throughput for latency of termination. As we show when evaluating the performance of our protocol, the average termination latency of blocks is at most a few seconds. In return, when running on non-dedicated virtual machines and network, in a single Amazon data-center, we demonstrate performance of up to 160K transactions per seconds. In a non-dedicated multi data-center settings, we obtain up to 30K transactions per second.

2 Related Work

Optimistic Consensus Two main methods were suggested for designing an optimistic consensus protocol: (i) satisfying safety from the nodes’ point of view [46, 53, 60] or (ii) satisfying safety only from the clients’ point of view [13, 19, 55]. In the first approach, in order to detect inconsistencies, nodes must continuously update other nodes with their state (the exception is [46] that uses randomization). In the second approach, nodes are allowed to be temporarily inconsistent with each other. Only when a client detects an inconsistency, e.g., by receiving inconsistent replies, it initiates a special recovery mechanism to restore the system’s consistency. Concerning blockchains, the first method ignores blockchain’s unique features that can be leveraged. In contrast, running the blockchain nodes as clients of an agreement service results in at least two communication steps protocol even in the “good cases”.

Blockchain Systems To circumvent FLP [45], most unpermissioned blockchain platforms such as Bitcoin [63], Ethereum [44] and Algorand [49] assume a synchronous network, signed messages and a Sybil prevention mechanism. Bitcoin and the current implementation of Ethereum rely on Proof-of-Work (PoW) while Algorand employs Proof-of-Stake (PoS). In both cases, the algorithm produces a new block in every time slot whose length depends on the worst case maximal network’s latency and block size. When the synchrony assumption is violated, safety might be violated as well, resulting in more than one version of the chain that exist concurrently, also known as a fork in the chain. To overcome forks the platform assumes an upper bound on the computation resources that are held by malicious nodes (less than 50%) and defines a deterministic rule in which a node always prefers the longest chain it knows about. Although with PoS the chain may fork with a probability that is less than one in a trillion, with PoW, forks may happen frequently. Hence, PoW and the above recovery mechanism impose another crucial drawback: a transaction is never permanent since a longer version may always emerge in the future (although with rapidly diminishing probability). Thus, clients cannot rely on a new block until it is deep enough in the chain, resulting in high latency even in the common case. PoS improves the performance w.r.t. PoW, but results in very complicated protocols.

Permissioned blockchain protocols typically assume a partially synchronous network while utilizing traditional BFT concepts. Such platforms run a more computationally efficient protocol than unpermissioned blockchains but require an a-priori PKI infrastructure. Traditional BFT solutions are not scalable in the number of participants [73] as their communication complexity grows quadratically in the number of nodes. Hence, such solutions focus on (i) sharding the execution’s roles between multiple layers, leaving the consensus to be run by a small set of nodes, and on (ii) designing optimized dedicated BFT consensus protocols. Known platforms like HLF [8, 5, 17] and R3 Corda [2] offer new models of layered computation and run the BFT-SMaRt [20] protocol, or variants of it, as an ordering service layer. HLF runs the three phases model of execute-order-validate. R3 Corda maintains a hashed directed acyclic graph named Hash-DAG (rather than a single chain), in which a transaction is stored only by those nodes who are affected by it. To ensure transactions’ validity, Corda offers a BFT-SMaRt based distributed notary service.
Platforms such as Chain Core [1] [30], Iroha [6], Symboint-Assembly [8] and Tendermint [23] offer new optimized BFT Consensus algorithms. Iroha, inspired by the original HLF (v0.6) architecture, runs the Sumeragi consensus protocol which is heavily inspired by BChain [40]. BChain is a chain-replication system in which $n$ nodes are linearly arranged and a transaction is moved among the nodes in a chain topology. Namely, each node normally receives a message only from its predecessor. Like FireLedger, BChain trades latency for throughput and it has the potential to achieve the best possible throughout [19] [50]. Unlike FireLedger, BChain’s latency is bounded by at least $n$ rounds. Symboint-Assembly implements its own variant of BFT-SMaRt. Tendermint implements an iterative variant of PBFT [29] designed by Buchman et al. [25]. Chain Core runs the Federated consensus protocol in which one node is the leader and $n$ are validators. This protocol is Byzantine resilient for $f < \frac{n}{3}$ only as long the leader is correct. Red Belly blockchain [7] offers both, a new computation model that balances the verification load among verifies nodes and the Democratic BFT consensus [36] that is able to scale the throughput with the number of proposers. Finally, HoneyBadger BFT (HBB) [60], [4] is a randomized protocol targeting blockchain. HBB circumvents FLP by randomization and is based on a probabilistic binary Byzantine consensus [61].

HotStuff [74] extends transactions’ finality to 3 rounds and employs signature aggregation [22] in order to obtain linear communication overhead. HotStuff requires all nodes to sign an asymmetric signature on each block in the optimistic case while in FireLedger this is done only by the proposer that generated the block. Since signing takes pure CPU time, fewer asymmetric signatures enable better throughput.

The notion of $k$-coherence was defined in [16]. FireLedger can be viewed as guaranteeing an adjustment of $(f+1)$-coherence for deterministic consensus [57], [41]. FireLedger guarantees the classical consensus properties only for decisions at depth greater than $f + 1$ and with respect to an external valid method [27], [36].

3 Preliminaries and Problem Statement

3.1 System Model

We consider an asynchronous fully connected environment consisting of $n$ nodes out of which at most $f < \frac{n}{3}$ may incur Byzantine failures [18], [60]. Asynchronous means that no upper bound on the messages’ transfer delays exists and nodes have no access to a global clock. While nodes may have access to local clocks, these clocks might not be synchronized with each other and may advance at different rates. Fully connected means that any two nodes are connected via a reliable link. Reliable means that a link does not lose, modify or duplicate a sent message. Notice that unreliable fair lossy links can be transformed into reliable ones using sequence numbering, retransmissions, and error detection codes [68]. A Byzantine failure means that a node might deviate arbitrarily w.r.t. its protocol code including, e.g., sending arbitrary messages, sending messages with different values to different nodes, or failing to send any or all messages. Yet, we assume that nodes cannot impersonate each other. A node suffering from a Byzantine failure at any point during its operation is called Byzantine; otherwise, it is said to be correct.

Following the FLP result [45], consensus is unsolvable in a truly asynchronous system. We circumvent FLP by enriching the system with the $\diamond Synch$ assumption [23]. $\diamond Synch$ means that after an unknown time $\tau$ there is an unknown upper bound $\delta$ on a message’s transfer delay. As in most Byzantine fault tolerance works [18], [29], [53], $\diamond Synch$ is only needed to ensure liveness, meaning that even under severe network delays safety is never violated. Finally, a node may sign a message by an unforgeable signature. We denote the signature of node $p$ on message $m$ by $sig_p(m)$. The implementation of the signature mechanism is done by a well known cryptographic technique, such as symmetric ciphers [42], RSA [67] or an elliptic curves digital signature (ECDS) [52].

3.2 Underlying Protocols

Solutions to the following fundamental distributed computing problems serve as building blocks for FireLedger.
Reliable Broadcast (RB) The reliable broadcast abstraction \cite{24} (denoted RB-Broadcast) ensures reliable message delivery in the presence of Byzantine failures. To utilize RB-Broadcast nodes may invoke two methods: RB-broadcast and RB-deliver. A correct node that wishes to broadcast a message \( m \) invokes RB-broadcast\( (m) \) while a node that expects to receive a message invokes RB-deliver. By a slight abuse of notation, we denote RB-deliver\( (m) \) the fact than an invocation of RB-deliver returned the message \( m \) and say that the invoking process has RB-delivered \( m \). The RB-Broadcast abstraction satisfies the following properties:

**RB-Validity:** If a correct node has RB-delivered a message \( m \) from a correct node \( p \), then \( p \) has invoked RB-broadcast\( (m) \).

**RB-Agreement:** If a correct node RB-delivers a message \( m \), then all correct nodes eventually RB-deliver \( m \).

**RB-Termination:** If a correct node invokes RB-broadcast\( (m) \), then all correct nodes eventually RB-deliver \( m \).

Atomic Broadcast (AB) The atomic broadcast abstraction \cite{37} is more restrictive than RB-Broadcast in the sense that in addition to the RB-Broadcast properties it requires the Order property as well.

**Atomic-Order:** All messages delivered by correct nodes are delivered in the same order by all correct nodes.

Multi-value Byzantine Consensus Most implementations of atomic broadcast rely on consensus protocols as sub-routines. In our context, Multi-value Byzantine Consensus (MVC) \cite{57} is a variant of the consensus problem in which a set of nodes, each potentially proposing its own value, must decide the same value in a decentralized network despite the presence of Byzantine failures. A solution to the MVC problem satisfies the following properties \cite{41}:

**MVC-Validity:** If all correct nodes have proposed the same value \( v \), then \( v \) must be decided.

**MVC-Agreement:** No two correct nodes decide differently.

**MVC-Termination:** Each correct node eventually decides.

Notice that according to MVC-Validity, any decision is valid when not all correct nodes propose the same value. Hence, it is not helpful when the system begins in a non-agreement state, which is common in real systems. To that end, an extended validity property was suggested in \cite{34} and is composed of three sub-properties:

**MVC1-Validity:** If all correct nodes have proposed the same value \( v \), then \( v \) must be decided.

**MVC2-Validity:** A decided value \( v \) was proposed by some node or \( v = nil \).

**MVC3-Validity:** No correct node decides a value that was proposed by only faulty nodes.

This validity property defines that if no correct node has suggested \( v \), then \( v \) cannot be decided. This definition states precisely what is the set of valid values in any possible run of the algorithm.

(Optimistic) Binary Byzantine Consensus The Binary Byzantine Consensus (BBC) is the simplest variant of MVC in which only two values are possible. The MVC-Validity, MVC-Agreement, and MVC-Termination properties naturally translate to their corresponding BBC-Validity, BBC-Agreement, and BBC-Termination properties. Interestingly, BBC cannot be solved in a weaker model than MVC. However, as only two values are possible, an Optimistic BBC (OBBC) is capable of achieving an agreement in a single communication step if a predefined set of conditions is met \cite{26,46,62}.

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3.3 Problem Statement

Blockchain algorithms require an external validity mechanism as sometimes even Byzantine nodes may propose legal values (or blocks) \[58\]. Therefore, the validity of a value may be defined by an external predefined method. The \textit{Validity Predicate-based Byzantine Consensus (VPBC)} \[27, 36\] abstraction captures this observation by defining the following validity property:

\textbf{VPBC-Validity:} A decided value satisfies an external predefined \texttt{VALID} method.

\textbf{VPBC-Agreement} and \textbf{VPBC-Termination} are the same as their MVC/BBC counterparts. That is, VPBC generalizes the classical definition of MVC.

In a blockchain, each block carries a glimpse to its creator knowledge of the system’s state. This glimpse is encapsulated in the hash that each block carries. In order to leverage the iterative nature of blockchains, we define a weaker model than \textit{VPBC} in which we denote each iteration with a round number \( r \). Next, we define the following per round notions:

\textbf{Tentative decision:} A decision of the protocol at a given node and round that might still be changed.

\textbf{Definite decision:} A decision of the protocol at a given node and round that will never change.

\( v^r_p \): A value that was decided or received by \( p \) in round \( r \) of the protocol. \( v^r \) denotes a value that was decided or received by some node in round \( r \).

\( d(v^r_p) \): Let \( r' \) be the current round of the protocol that node \( p \) runs. For a given \( v^r_p \) (possible tentative), we denote its \textit{depth} as \( d(v^r_p) = r' - r \).

\textbf{Definition 3.3.1} (\textit{Blockchain Based Finality Consensus (BBFC)}). Let \texttt{valid} be a predefined method as in VPBC and let \( \rho \) be a predefined fixed constant. The \( \rho \)-\textit{Blockchain Based Finality Consensus (BBFC(\( \rho \)))} abstraction defines the following properties:

\textbf{BBFC-Validity:} A decided (possible tentative) value \( v \) satisfies the \texttt{VALID} method.

\textbf{BBFC-Agreement:} For any two correct nodes \( p, q \), let \( v^r_p, v^r_q \) be their decided value in round \( r \). If \( d(v^r_p), d(v^r_q) > \rho \) then \( v^r_p = v^r_q \).

\textbf{BBFC-Termination:} Every round eventually terminates.

\textbf{BBFC-Finality:} In every round \( r' > r + \rho \), \( v^r_p \) is definite.

BBFC guarantees the VPBC’s properties only for decisions at depth greater than \( \rho \). With blockchains, as every block contains an authentication data regarding its predecessors, it provides a cryptographic summary of its creator history. This information may assist in detecting failures without the necessity of sending more information. In addition, blocks are continuously added to the chain. Thus, a block eventually becomes deep enough such that it satisfies the standard VPBC properties. Let us note that the BBFC-Agreement property is similar to the notion of \textit{common prefix} in \[48\].

In a blockchain setting, clients of the system submit \textit{transactions} to the nodes, and the decisions values are blocks, each consisting of zero or more transactions previously submitted by clients. In case the \texttt{VALID} method may accept empty blocks, we would like to prevent trivial implementations in which every node locally generates empty blocks continuously. Obviously, the throughput of such a protocol would be 0, and thus it would be considered useless. Yet, in order to prove that a protocol does not unintentionally suffer from such a behavior even under Byzantine failures we add the following requirement:

\textbf{Non-Triviality:} If clients repeatedly submit transactions to the system, then the nodes repeatedly decide definitively non-empty blocks.
4 Weak Reliable Broadcast

4.1 Overview

The Weak Reliable Broadcast (WRB) abstraction serves as FireLedger’s main message dissemination mechanism. WRB offers weaker agreement guarantees than RB-Broadcast [24]. In general, WRB ensures that the nodes agree on (i) the sender’s identity and (ii) whether to deliver a message at all, rather than the content of the message\(^1\). WRB is associated with the WRB-broadcast and WRB-deliver methods. A node that wishes to disseminate a message\( m \) invokes WRB-broadcast\( (m) \). If a node expects to receive a message\( m \) from\( k \) through this mechanism it invokes WRB-deliver\( (k) \). WRB-deliver\( (k) \) returns a message\( m \) where\( m \) is the received message. If the nodes were not able to deliver\( k \)’s message, WRB-deliver\( (k) \) returns nil. Formally, WRB satisfies the following properties:

WRB-Validity: If a correct node WRB-delivered\( (k) \)\( m \neq \text{nil} \), then\( k \) has invoked WRB-broadcast\( (m) \).

WRB-Agreement: If two correct nodes\( p, q \) WRB-delivered\( (k) \)\( m_p, m_q \) respectively from\( k \), then either\( m_p = m_q = \text{nil} \) or\( m_p \neq \text{nil} \land m_q \neq \text{nil} \).

WRB-Termination: If a correct node\( p \) WRB-deliver\( (k) \)\( m \) from\( k \), then every correct node that is trying to WRB-deliver\( (k) \) eventually WRB-deliver\( (k) \) some message\( m' \) from\( k \).

WRB-Non-Triviality: If a correct node\( k \) repeatedly invokes WRB-broadcast\( (m) \) then eventually all correct nodes will WRB-deliver\( (k) \)\( m \).

4.2 Implementing WRB

The pseudo-code implementation of WRB is listed in Algorithm 1. To WRB-broadcast a message, a node simply broadcasts it to everyone (line 3). When\( p \) invokes WRB-deliver\( (k) \) it performs the following:

- It waits for at most\( \text{timer} \) to receive\( k \)’s message (line 7).
- If such a valid message has been received, then\( p \) votes to deliver it using an OBBC protocol. Else,\( p \) votes against delivering\( k \)’s message (lines 8-13). Recall that if no node has proposed 0, then OBBC ends in a single communication step.
- If the decision is not to deliver (OBBC returned 0),\( p \) returns\( \text{nil} \) and increases the timer (lines 14-17).
- If it is decided to deliver the message (OBBC returned 1) and the message has already been received by\( p \), then\( p \) adjusts the timer and returns\( m \) (lines 18-21).
- Else, OBBC decided 1, meaning that at least one correct node received\( k \)’s message and voted for its acceptance. Thus,\( p \) moves to a pull phase and pulls\( k \)’s message from the nodes who did receive it. When\( p \) eventually receives a valid message,\( p \) returns it (lines 22-24).
- Upon receiving\( q \)’s request for\( k \)’s message, if\( p \) has\( k \)’s message\( m \), it sends\( (m, \text{sig}_k(m)) \) back to\( q \) (lines 25-27).

To ensure liveness, the\( \text{timer} \) is increased each time\( p \) does not receive the message (line 14). To avoid having too long timers for too long,\( \text{timer} \) is adjusted downward when a message is received by\( p \) (line 19). The exact details of such adjustments are beyond the scope of this paper, but see for example [29].

In a typical implementation of OBBC each node broadcasts its vote [20, 66]. Then if a node receives enough votes for the same value\( v \) to safely decide\( v \) after this single communication step, it decides\( v \) and returns. We present our own OBBC protocol in Appendix A.

\(^1\) To the best of our knowledge, we are the first to discuss WRB. Bracha’s approach [24] can be viewed as the opposite, first agree on the content of the message (consistent broadcast), and then agree whether it should be delivered.
Algorithm 1: Weak Reliable Broadcast - code for p

```
Procedure WRB-broadcast(m)
    timer ← τ;
    broadcast (m, sig_k(m)); /* push phase */

Procedure WRB-deliver(k)
    wait until a valid (m, sig_k(m)) has been received or timer has expired;
    if a valid (m, sig_k(m)) has been received then
        d ← OBBC.propose(1);
    /* If no node has proposed 0, then OBBC.propose ends in a single communication step */
    else
        d ← OBBC.propose(0);
        increase timer;
        if d = 0 then
            return nil;
        end
        broadcast(REQ,k);
        wait until a valid (m', sig_k(m')) has been received;
        return m';
        upon receiving (REQ,k) from q∧ a valid (m, sig_k(m)) has been received do
        send(m, sig_k(m)) to q;
    end
```

4.3 WRB Correctness Proof

Lemma 4.3.1 (WRB-Validity). If a correct node p WRB-delivered(k) a message m ≠ nil, then k has invoked WRB-broadcast(m).

Proof. The system model assumes reliable channels. Also, each node signs the messages it WRB-broadcasts. Hence, if a message m is received and pass validation as being sent by k, then k has indeed WRB-broadcast(m). □

Lemma 4.3.2 (WRB-Agreement). If two correct nodes p, q WRB-delivered(k) m_p, m_q respectively from k, then m_p, m_q ≠ nil. Namely, all correct nodes either deliver a message different than nil from k or they all deliver nil.

Proof. Let p, q be two correct nodes that have WRB-delivered(k) m_p, m_q respectively from k. Assume b.w.o.c and w.l.o.g that m_p ≠ nil while m_q = nil. A correct node may return nil only if OBBC.propose returns 0. Respectively, a correct node returns m ≠ nil only if OBBC.propose returned 1. This is in contradiction to BBC-Agreement. Hence, either m_p, m_q ≠ nil or both p and q return nil. □

Lemma 4.3.3 (WRB-Termination). If a correct node p WRB-deliver(k) m from k, then every correct node that is trying to WRB-deliver(k) eventually WRB-deliver(k) some message m’ from k.

Proof. A node p waits for k’s message at most timer time units. Hence p eventually invokes OBBC.propose. By BBC-Termination, no correct node is blocked forever on OBBC.propose. If OBBC.propose has returned 0 then p returns nil. Else, if k’s message has been received by p it returns m. Otherwise, p asks from other nodes for k’s message. By BBC-Validity, if OBBC.propose returns a value v then at least one correct node q proposed v. Thus, if OBBC.propose returned 1 it means that at least one correct node q has proposed 1 and by the algorithm’s code q does so if it has received k’s message m. Hence, when q receives p’s request it sends back m. As q is correct p receives and returns m. □
Lemma 4.3.4 (WRB-Non-Triviality). If a correct node $k$ repeatedly invokes $WRB$-$broadcast(m)$ then eventually all correct nodes will $WRB$-$deliver(k)m$.

Proof. By the algorithm’s code, on any unsuccessful delivery the timer increases. Let $k$ be a correct node that repeatedly invokes $WRB$-$broadcast(m)$. Following the $\&$Synch assumption and the incremental timer method [29], eventually after an unknown time $\tau$ the timer of all nodes is longer than the unknown upper bound $\delta$ on a message’s transfer delay. Hence, every correct node receives $k$’s message and invokes $OBBC.propose(1)$. By $BBC$-$Agreement$ if all correct nodes propose 1 then 1 has to be decided. Hence, every correct node eventually $WRB$-$deliver(k)m$. $\square$

By the above four lemmas we have:

Theorem 4.3.1. The protocol listed in Algorithm 1 solves WRB.

Note that WRB’s communication costs depend on the success of OBBC. If OBBC decides in a single communication step, then WRB-deliver also terminates in a single communication step.

5 The FireLedger Protocol

FireLedger implements the $BBFC(\rho)$ abstraction with $\rho = f + 1$. This section focuses on the algorithmic aspects of FireLedger and its correctness, while actual implementation considerations and performance are discussed in Section 6 below.

We present FireLedger in a didactic way: We first show a two-phased crash fault tolerant (CFT) ordering protocol based on WRB. We then improve it to a single-phased protocol. Finally, we extend the protocol to tolerate Byzantine failures. The pseudo-code of the protocol appears in Algorithm 2 and 3. Regularly numbered lines correspond to the CFT aspects of the protocol, while lines prefixed with ‘b’ are the additions to handle Byzantine failures.

5.1 Simplified CFT FireLedger

As mentioned in Section 4, WRB supports an all or nothing delivery that is blind to the message’s content. When there are no Byzantine failures, a node never sends different messages to different nodes. Hence, a simple two phase blockchain protocol would be a round based design in which a deterministically selected leader disseminates its block proposal to all nodes using WRB. In case all nodes deliver the proposed block, then this becomes the next block.

Given the continuous iterative nature of blockchains, we may improve the algorithm’s round latency to an amortized single phase. For this, we piggyback the $(r + 1)^{th}$ block on top of the first message that $WRB$-$deliver$ sends when trying to deliver the $r^{th}$ block. We support this by augmenting the $WRB$-$deliver$ method to receive two parameters, $WRB$-$deliver(k, pgd)$, such that $pgd$ is the potential piggybacked data (can be $nil$). The OBBC protocol is also augmented to receive $pgd$ and piggyback it on the first message it broadcasts. Recall that we assume an OBBC protocol that always starts by having each node broadcast its vote. Hence in the augmented protocol, each node that starts the OBBC protocol broadcasts its vote alongside the piggybacked $pgd$ message, which is made available to the calling code together with the decision value.

The details appear in Algorithm 2. Specifically, on each round $r$ the algorithm performs the following:

- If $p_i$ is $(r + 1)$’s proposer, it prepares a new block (lines 12-14). To ensure liveness, if in the previous iteration WRB has failed to deliver a message, then $r$’s proposer also prepares a block and WRB-broadcast it (lines 8-11).

- Meanwhile, all nodes are trying to $WRB$-$deliver r$’s block (line 15). Note that the $(r + 1)^{th}$ proposer piggybacks the next block on top of this message.
Algorithm 2: FireLedger – code for $p_i$
The lines that start with ‘b’ depict the BFT additions

```plaintext
1  $r_i \leftarrow 0$;
2  $proposer_{r_i} \leftarrow p_0$;
3  $full\_mode \leftarrow true$;
4  while true do
5    $b \leftarrow nil$;
6      while $proposer_{r_i}$'s block was tentatively decided in the last $f$ rounds do
7        $proposer_{r_i} \leftarrow (proposer_{r_i} + 1) \mod n$;
8      end
9      if $i = proposer_{r_i} \land full\_mode = true$ then
10         /* executed if nil has been WRB-delivered in the last iteration */
11         $b \leftarrow prepared\_block$;
12         WRB-broadcast($b$);
13         $b \leftarrow nil$;
14      end
15      if $(proposer_{r_i} + 1) \mod n = i$ then
16         $b \leftarrow prepared\_block$;
17      end
18      $b_{r_i} \leftarrow WRB\_deliver(proposer_{r_i}, b)$;
19      if $b_{r_i} = nil$ then
20         $full\_mode \leftarrow true$;
21         $proposer_{r_i} \leftarrow (proposer_{r_i} + 1) \mod n$;
22         continue;
23      end
24      full\_mode \leftarrow false;
25      if $b_{r_i}$ is not valid then
26         /* validating the block hash against the previous block */
27         $proof \leftarrow (b_{r_i}, \text{sig}_{proposer_{r_i}}(b_{r_i}), b_{r_i-1}, \text{sig}_{proposer_{r_i-1}}(b_{r_i-1}))$;
28         RB-broadcast($proof$);
29         invoke RECOVERY($r_i, proof$);
30         continue;
31      end
32      append $b_{r_i}$ to the chain;
33      decide $b_r^{(r+2)}$;
34      $proposer_{r_i} \leftarrow (proposer_{r_i} + 1) \mod n$;
35      $r_i \leftarrow r_i + 1$;
36  end
37  upon RB-deliver a valid $proof \leftarrow (b^r, \text{sig}_{proposer_{r}}, (b^{r-1}, \text{sig}_{proposer_{r-1}}(b^{r-1})))$ do
38    invoke RECOVERY($r, proof$);
39  end
```

- If a block $b^r \neq nil$ has been delivered, then $b^r$ is appended to the chain (line 22) and the protocol continues to round $r + 1$ (line 23-25).
- Else, all correct nodes switch proposer and continue to the next try (lines 16-20).

We prove the correctness of the full BFT protocol (Section 5.3). Note that due to WRB and the piggybacking method, algorithm 2 establishes a single-phased protocol as long there are no Byzantine failures, and the failure pattern matches the specific OBBC optimistic pattern.

Figure 1 presents normal case operation. Each optimistic period starts with the current proposer broadcasting its block. Then, on every round, each node broadcasts a single message (as the first OBBC message), except the next proposer that piggybacks the next block on top of that message.
Algorithm 3: Recovery Procedure – code for $p_i$

```
(1) Procedure RECOVERY($r$, proof)
(2)  versions_r ← {};
(3)  if $r_i < r - 1$ then
(4)    v ← empty_version;
(5)  else
(6)    v ← ($b^{r-(f+1)}$, sig_proposer$_{r_i-(f+1)}$($b^{r-(f+1)}$), ..., $b^{r-1}$, sig_proposer$_{r_i-1}$($b^{r-1}$), ..., $b^{r_i}$, sig_proposer$_{r_i}$($b^{r_i}$));
(7) end
(8) Atomic-broadcast(v);
(9) repeat
(10)  Atomic-deliver $v_j$ from $p_j$;
(11)  if $v_j$ is valid then
(12)    versions_r ← versions_r ∪ {$v_j$};
(13)  end
(14) until |versions_r| = n - f;
(15) $v'$ ← the first received among \{$v_j \in$ versions_r | $r_j = \max\{r_k | v_k \in$ versions_r ∧ ($b^k$, sig_proposer$_k$($b^k$)) ∈ $v_k$\};
(16) adopt $v'$ and update $r_i$ and proposer$_{r_i}$;
(17) full_mode ← true;
```

5.2 Full BFT FireLedger

In order to extend the basic FireLedger to handle Byzantine failures, FireLedger utilizes the fact that there is at least one correct node $p_c$ in every $f + 1$ different proposers. Since $p_c$ is correct, when $p_c$'s block is WRB-delivered, all correct nodes receive the very same block, including the hash to its predecessor block. When a correct node detects a chain inconsistency (due to the hash that each block carries), it initiates a traditional BFT based recovery procedure. At the end of the recovery phase, all correct nodes synchronize their chain to the same single valid version. To enable the recovery procedure, FireLedger maintains the following invariant:

**Invariant 1.** A node $p$ proceeds to the next round of the algorithm only if it knows that at least $f + 1$ correct nodes will eventually proceed as well.

An immediate consequence of Invariant 1 is that if a block $b^r$ is at depth $f + 1$ in $p$'s local chain, then there are at least $f + 1$ correct nodes for which $b^r$ is at depth of at least $f + 1$ in their local chains. In principle, preserving Invariant 1 requires waiting to verify that at least $f$ other correct nodes are moving to the next round. Yet, since the communication pattern we described above already includes an all-to-all message exchange in each round (while executing OBBC), it serves as an implicit acknowledgement, so we still have a single-phased algorithm when the optimistic assumptions hold.

Recall that line numbers prefixed by ‘b’ Algorithm 2 describe the additional actions to accommodate Byzantine failures. Specifically:

- First $p_i$ finds, by a pre-defined order, a prosper that has not successfully proposed a block in the last $f + 1$ rounds (lines b1- b3).
- If $p_i$ detects an inconsistency in the chain, it announces a recovery procedure using reliable broadcast (lines b4- b10).
- Upon receiving a valid announcement of inconsistency, $p_i$ initiates the recovery procedure (lines b12- b14). Note that the announcement validation is done against digital signatures and the blocks’ hashes.

The recovery procedure installs agreement among all correct nodes regarding the longest possible blockchain prefix as detailed in Algorithm 3. Executing RECOVERY by $p_i$ involves:
• \( p_i \) proposes, using \( \text{Atomic-broadcast} \), a valid version of the \( f \) blocks that are in disagreement (excluding \( b^r \) itself) followed by all the newer blocks it knows about (lines 2–8).

• Then \( p_i \) collects \( n - f \) valid versions (including empty ones) and adopts the first longest agreed prefix of the blockchain (lines 9–18).

Finally, as the recovery procedure may alter only the last \( f + 1 \) blocks, the node decides on the block which is in depth of \( f + 2 \) (line 11).

### 5.3 Correctness Proof

**Lemma 5.3.1** (BBFC\((f + 1)\)-\text{Validity}). A decided (possible tentative) value \( v \) satisfies the \text{VALID} method.

*Proof.* From Algorithm 2’s code, a value is appended to the blockchain only if it satisfies the \text{VALID} method.

**Lemma 5.3.2.** Every \( f + 1 \) consecutive decided blocks were proposed by \( f + 1 \) different nodes.

*Proof.* By Algorithm 2’s code (lines 11–13), if the current proposer has successfully proposed a block in the last \( f + 1 \) rounds, then its role is switched to a new proposer.

**Lemma 5.3.3.** At the recovery procedure’s end, all correct nodes adopt the same version.

*Proof.* By the \text{Atomic-Order} property of the atomic broadcast primitive all correct nodes receive the same versions in the same order. Hence, applying the deterministic rule described in Algorithm 2 results in an agreement among the correct nodes.

**Lemma 5.3.4.** For any two correct nodes \( p, q \), let \( b^r_p, b^r_q \) be their decided value in round \( r \). If \( d(b^r_p) > f \land d(b^r_q) > f \) then \( b^r_p = b^r_q \).

*Proof.* Let \( p, q \) be two correct nodes and let \( b^r_p, b^r_q \) be their decided value in round \( r \) such that \( d(b^r_p) > f \land d(b^r_q) > f \) and assume b.w.o.c. that \( b^r_p \neq b^r_q \).

Let \( p_r \) be \( r \)’s proposer. If \( p_r \) is correct, then by \text{WRB-Agreement} and \( p_r \)’s correctness all correct nodes have received the very same message from \( p_r \) (as in Reliable Broadcast). Hence, the case of \( b^r_p \neq b^r_q \) may occur only as a result of the recovery procedure invocation in round \( r' \in \{r + 1, ..., r + f\} \). By Lemma 5.3.3 at the end of the recovery procedure all correct nodes agree on the same version. Hence, \( b^r_p = b^r_q \), contrary to the assumption.
Suppose that \( p_r \) is Byzantine and has disseminated different blocks to \( p \) and \( q \). Then, by Lemma \([5.3.2]\) and the fact that the system model assumes at most \( f \) faulty nodes there is at least one block \( b^{r'} \in \{ b^{r+1}, \ldots, b^{r+f} \} \) that was proposed by correct proposer \( p_r \). Due to \( p_r \)'s correctness it disseminates the same block to all, including the associated authentication data. Assume w.l.o.g. that \( p_r \) has received from \( p \) the same version as \( p \), then following Algorithm \([2]\) lines \([1]\)-\([10]\) \( q \) receives an invalid block from \( p_r \) and invokes the recovery procedure. As \( q \) is correct and by \( \text{RB-Termination} \) \( p \) RB-delivers \( q \)'s proof of \( b_q^{r'} \) invalidity and triggers the recovery procedure. Following Lemma \([5.3.3]\) at the end of the recovery procedure all correct nodes adopt the very same version. Hence, \( b_p^r = b_q^r \). A contradiction.

\[ \Box \]

**Lemma 5.3.5** (BBFC\((f + 1)\)-Agreement). For any two correct nodes \( p, q \), let \( b_p^r, b_q^r \) be their decided value in round \( r \). If \( d(b_p^r) > f + 1 \land d(b_q^r) > f + 1 \) then \( b_p^r = b_q^r \).

**Proof.** Follows directly from Lemma \([5.3.4]\)

**Definition 5.3.1.** Let \( b_p^r, b_p^{r'} \) be two decided (possibly tentative) blocks of \( p \) such that \( r' > r \). \( b_p^{r'} \) is valid with respect to \( b_p^r \) if the sub-chain \( [b_p^r, b_p^{r+1}, \ldots, b_p^{r'}] \) satisfies the predefined VALID method. Similarly, a sub-chain \( [b_p^{r'}, b_p^{r'+1}, \ldots, b_p^{r''}] \) is valid with respect to \( b_p^{r''-k} \), for some \( k \leq r \), if the sub-chain \( [b_p^{r''-k}, \ldots, b_p^{r'}, \ldots, b_p^r] \) satisfies the predefined VALID method and each \( f + 1 \) consecutive blocks are proposed by \( f + 1 \) different proposers.

**Lemma 5.3.6.** If during the recovery procedure for round \( r \), a correct node \( p \) receives a version \( v \) from correct node \( q \), then \( v \) is valid with respect to \( b_p^{r-(f+2)} \).

**Proof.** If the received version is an empty one, it is trivially valid with respect to \( b_p^{r-(f+2)} \). Else, by Lemma \([5.3.4]\) all correct nodes agree on \( b_p^{r-(f+2)} \). As \( q \) is correct, by Lemma \([5.3.1]\) \( q \) appends only valid blocks to its blockchain. Hence, \( q \)'s version (that starts with \( b_q^{r-(f+1)} \)) is valid with respect to \( b_q^{r-(f+2)} \) which is identical to \( b_p^{r-(f+2)} \).

**Definition 5.3.2.** Let \( r_g \) be the most advanced round of the algorithm that any correct node runs. We define the group of nodes whose current round is \( r_g \) by \( \text{front} = \{ p | r_p \in \{ r_g - 1, r_g \} \} \). When Invariant \([1]\) is kept, there are at least \( f + 1 \) correct nodes in \( \text{front} \).

**Lemma 5.3.7.** While Invariant \([1]\) is kept, a correct node executing the recovery procedure receives \( n - f \) valid versions and at least one of them was received from a node in \( \text{front} \).

**Proof.** By \( \text{RB-Termination} \) if a correct node \( p \) detects an invalid block and invokes the recovery procedure, eventually every correct node will receive \( p \)'s proof and will invoke the recovery procedure. Following the system model, the ratio between \( n \) and \( f \) and Lemma \([5.3.6]\) a correct node does not get blocked while waiting for \( n - f \) valid versions (part of whom may be empty). By Invariant \([1]\) \( p \) receives at least one version from a correct node in \( \text{front} \).

**Lemma 5.3.8** (BBFC\((f + 1)\)-Termination). Every round eventually terminates.

**Proof.** Let \( r \) be a round of the algorithm. As long WRB-deliver has returned \( \text{nil} \), a correct proposer will propose its next block. Thus, by \( \text{WRB-Termination}(1), \text{WRB-Termination}(2) \) and the algorithm’s iterative nature every correct node eventually succeeds to \( \text{WRB-deliver} \ b^r \neq \text{nil} \).

If \( b^r \) is valid and the recovery procedure has not been invoked, then naturally nothing prevents the round from terminating. If the recovery procedure has been invoked, then following Lemma \([5.3.7]\) every correct node eventually receives \( n - f \) valid versions and finishes the recovery procedure; in the worst case, such a node stays in round \( r - 1 \). The system model assumes at most \( f \) faulty nodes, which means that the above can repeat itself at most \( f \) times in a row. Hence, after at most \( f \) consecutive invocations of the recovery procedure, a correct proposer proposes \( b^r \), ensuring that \( r \) terminates in all correct nodes.

**Definition 5.3.3.** Let \( r \) be a round of the algorithm, we define by \( \text{front}_r = \{ p | r_p > r + f + 1 \} \) the group of nodes whose current round is greater by at least \( f + 1 \) rounds than \( r \).
Figure 2: If FireLedger would implement BBFC($f$) instead of BBFC($f + 1$), a node $q$ may replace a definite block. While recovering, due to contention on $b_q^r$, the nodes may adopt a prefix that was suggested by $p \in \text{front}$ while $p$ is still in round $r$.

Lemma 5.3.9 (BBFC($f + 1$)-Finality). In every round $r' > r + f + 1$, $v_p^{r'}$ is definite.

Proof. Following Algorithm 2's code, after a valid block is WRB-delivered, the only way it can be replaced is by an invocation of the recovery procedure. By Algorithm 2's code and Lemma 5.3.7 if the recovery procedure has been invoked regarding round $r$, then $p$ adopts a prefix version that was suggested by $q \in \text{front}$. Obviously, $\text{front} \subseteq \text{front}_r$. By Lemma 5.3.5, $\forall q, p \in \text{front}_r$, $b_q^r = b_p^r$. In other words, $b_p^r$ is definite.

Note that the ratio between $n$ and $f$ as well as the version’s validation test ensure that no Byzantine node is able to propose a version that does not include $b_q^r$.

Notice that due to asynchrony, it is possible that some node $q \in \text{front}$ advances to round $r + 1$ while another node $p \in \text{front}$ invokes the recovery procedure for round $r + 1$. In such a case, $q$ may adopt a new blockchain prefix which was proposed by $p$ and thus replace $b_q^{r-(f+1)}$. This edge case, illustrated in Figure 2, is the reason why FireLedger implements BBFC($f + 1$) rather than BBFC($f$).

The next theorem follows from Lemmas 5.3.1, 5.3.5, 5.3.8 and 5.3.9:

Theorem 5.3.1. The protocol listed in Algorithm 2 solves BBFC($f + 1$).

Lemma 5.3.10 (Atomic Order). All blocks delivered by correct nodes using Algorithm 2 are delivered in the same order.

Proof. Follows directly from Lemma 5.3.5 and the iterative nature of the algorithm.

By Lemma 5.3.10 and Theorem 5.3.1 we have:

Theorem 5.3.2. The protocol listed in Algorithm 2 imposes a total order of all blocks.

Theorem 5.3.3. The protocol listed in Algorithm 2 satisfies Non-Triviality.

The proof of Theorem 5.3.3 follows directly from Lemmas 5.3.8 and 5.3.9 as well as the observation that correct nodes propose non-empty blocks whenever they hold clients’ transactions that have not been included in any previously decided block.

5.4 Theoretical Bounds and Performance

Table 1 summarizes the performance of FireLedger in each of its three modes. In case of no failures and synchronized network, FireLedger performs in the OBBC of WRB-deliver a single all-to-all communication step as well as one digital signature operation. Further, only one node broadcasts more than one bit. This can be obtained since the latency for a definitive decision can be up to $f + 1$ rounds. In case a message is not received by WRB-deliver due to timing, omission or benign failures, the algorithm runs the non optimistic phase of OBBC as well as two more communication steps (one-to-all and one-to-one) in which a node asks for the missed message from the nodes who did receive it. Also, the amount of digital signature operations depends on the specific OBBC implementation that is used by WRB-deliver. Finally, in case of Byzantine
failures, FireLedger runs the recovery procedure which depends on the Reliable Broadcast and the Atomic Broadcast implementations.

Note that one can trade additional communication steps for fewer digital signature operations by using signature-free implementation of FireLedger’s base protocols, e.g., [61, 46, 35]. As FireLedger is inherently not signature-free, a node may authenticate up to $n - f$ received prefix versions. Yet, when Byzantine failures manifest, a node does not lose blocks so the latency in terms of rounds remains the same.

6 Implementation

6.1 FireLedger’s Implementation

We now describe FireLedger’s implementation components as well as a few necessary performance optimizations.

6.1.1 Optimizations

Separating Headers and Blocks FireLedger’s protocol enables to easily separate the data path from the consensus path, such that only block headers need to pass through the consensus layer while the block itself is being sent asynchronously in the background.

In practice, a node $p$ broadcasts a block as soon as the block is ready. On $p$’s next proposing round, $p$ WRB-broadcasts a header of a previously sent block. Respectively, upon WRB-delivering a header, if $p$ did not receive the block, it votes against delivering it (See Algorithm 1, lines 8–13). In addition, if a decision was made to deliver a header, but the block itself has not been received by $p$, then $p$ has to retrieve the block from a correct node $q$ that has it. Such a $q$ exists because the decision to deliver is done only if at least one correct node has voted in favor of delivering, which means that $q$ has the block.

Dynamically Tuning the Timeout To adjust the timer to the current network delays status, we dynamically adjust its value based on the exponential moving average (EMA) of the message delays over the last $N$ rounds. Namely, denote by $d_k, timer_k$ the delay of a message and the timer of round $k$ respectively.

Then for every round $r$

$$timer_r = \frac{2}{N + 1} \cdot d_{r-1} + timer_{r-2} \cdot (1 - \frac{2}{N + 1}).$$

A formal discussion of the above tuning model is out of the scope of this paper.
**Benign FD** FireLedger’s algorithm enables implementing a simple benign failure detector (FD) such that a crashed node will not cause an unrestricted increase in the timer value. Largely speaking, every node $p$ maintains a suspected list of the $f$ nodes to which $p$ has waited the most and above a predefined threshold. For every node $q$ in that list, on a WRB-deliver, $p$ does not wait for $q$’s message but rather immediately votes against delivering. By the ratio between $n$ and $f$ there is at least one correct node $c$ that is not suspected by any correct node and thus the algorithm’s liveness still holds. Despite the above, if $c$ is one of the last $f$ proposers it would not be able to suggest a new block. Hence, the suspected list is invalidate every time FireLedger is skipping a node that is in the last $f$ proposers (see Algorithm 2, lines 1–3). Also, if a Byzantine activity was detected, to avoid considering more than $f$ nodes as faulty, we invalidate the suspected list.

**Consecutive Byzantine Proposers** Recall that nodes are chosen by default to serve as the initial proposers of each block in a round-robin manner. Hence, as presented so far, FireLedger’s performance might suffer if multiple Byzantine nodes are placed in consecutive places in this round-robin order. This is easily solvable by periodically changing the round-robin order to be a pseudo-randomly selected permutation of the set of nodes where the seed is unknown a-priory to the (Byzantine) nodes. For example, it can be the result of a *verifiable random function* (VRF) \[49\] whose seed is a given block’s hash value. In case the adversary is static, this can be done only once, or if the performance suddenly drops due to the activity of consecutive Byzantine nodes. To overcome a dynamic adversary whose minimal transfer time is equal to the time required to agree on $k$ blocks, we can simply invoke the above every $k$ blocks.

6.1.2 FireLedger’s Instance Implementation

Figure 3 depicts the main components of a FireLedger’s instance. Using FireLedger’s API one can feed the $TX$ pool with a new write request. The *main thread* creates a new block and WRB-broadcasts it in its turn. In addition, the *main thread* tries to WRB-deliver blocks relying on OBBC. If it succeeds, the block is added to the *Blockchain*. Meanwhile, the *panic thread* waits for a panic message. When such a message is *Atomic-delivered*, the *panic thread* interrupts the *main thread* which as a result invokes the *recovery* procedure. FireLedger is implemented in Java and the communication infrastructure uses gRPC, excluding BFT-SMaRt which has its own communication infrastructure. *Atomic Broadcast* is natively implemented on top of *BFT-SMaRt* whereas *OBBC* uses *BFT-SMaRt* only as the fallback mechanism when agreement cannot be reached through the optimistic fast path (which relies on gRPC).

6.2 FLO – FireLedger Orchestrator

While FireLedger assumes partial synchrony, its rotating leader pattern imposes synchronization in the sense that a node may propose a value only on its turn, making FireLedger’s throughput bounded by the actual network’s latency. To ameliorate this problem, we introduce another level of abstraction, named *workers*, by which each node runs multiple instances of FireLedger and uses them as a blockchain based ordering service. The use of workers brings two benefits: (i) workers behave asynchronously to each other which compensates for the above synchrony effect of FireLedger and (ii) while a worker waits for a message, other workers are able to run, resulting in better CPU utilization. To preserve the overall total ordering property of FireLedger, a node must collect the results from its workers in a pre-defined order, e.g., round robin. This requirement may imposes higher latency when the system is heavily loaded because even if a single worker faces the non-optimistic case, it delays all other workers from delivering their blocks to the node.

Figure 4 depicts an overview of a FLO node. Upon receiving a write request, the *client manager* directs the request to the least loaded worker. When a read request is received, the *client manager* tries to read the answer from the relevant worker. Only if the answer was already definitely decided and its block can be delivered in the pre-defined order, the node returns the answer.\[2\]

\[2\] This is similar to the *deterministic merge* idea of [14].
7 Performance Evaluation

Our evaluation studied the following questions:

- Is FLO/FireLedger CPU bounded or network bounded?
- How Table 2’s values affect FLO/FireLedger performance?
- How the node distribution method affects FLO/FireLedger performance?
- How does FLO/FireLedger handle failures?

Deployment Specification Our setup for most measurements includes \( n \) nodes running on \( n \) identical VMs with the following specification: \( \text{m5.xlarge} \) with 4 vCPUs of Intel Xeon Platinum 8175 2.5 GHz processor, 16 GiB memory and up to 10 Gbps network links (Section 7.6 uses a stronger configuration as detailed there). In all measurements, transactions are randomly generated.
| parameter     | range           | units |
|---------------|-----------------|-------|
| cluster size  | \( n \in \{4,7,10\} \) | -     |
| workers       | \( 1 \leq \omega \leq 10 \) | -     |
| transaction size | \( \sigma \in \{512,1K,4K\} \) | Byte  |
| batch size    | \( \beta \in \{10,100,1000\} \) | Transaction |

Table 2: The default evaluation’s parameters: The first row presents FLO’s cluster size. The system model assumes \( f < \frac{n}{3} \), hence, \( n \in \{4,7,10\} \) imposes \( f \in \{1,2,3\} \) respectively.

### 7.1 Signature Generation

In FLO/FireLedger, as long as the optimistic assumptions hold, a proposer signs its block only once and any other node is verifying the signature only once. Hence, the maximal signature rate serves as an upper bound on the potential throughput of FireLedger.

To understand whether FLO/FireLedger is CPU bounded or I/O bounded we start by presenting an evaluation of the signatures generation rate (sps) which is typically the CPU most intensive task. We use ECDSA signatures with the secp256k1 curve. When signing a block, all the block’s transactions are hashed and the result is signed alongside the block header. We vary the \( \omega, \beta \) and \( \sigma \) values in the ranges described in Table 2.

Denote by \( t_{\text{hash}} \) the time that takes to hash a single byte. For a block consisting of \( \beta \) transactions of \( \sigma \) bytes each, the block’s signing time, \( t_{\text{sign}} \), is expected to be

\[
t_{\text{sign}} = \beta \cdot \sigma \cdot t_{\text{hash}} + C
\]

where \( C \) is a constant representing the time that takes to sign the fixed sized block header.

For each configuration we run the benchmark for 1 minute. Figure 5 depicts the benchmark results. As expected, with small blocks the sps is higher than with larger ones. Also, as our machines have 4 vCPUs, increasing \( \omega \) beyond 4 has a minor effect if any.

Denote by \( tps \) the transactions per second rate. The following bound must hold:

\[
tps \leq \text{sps} \cdot \beta.
\]

As seen below, the performance of FLO is not limited by the sps rate.

### 7.2 FLO Cluster in a Single Data-Center

We deployed a FLO cluster where all machines reside in the same data center. To test the system in extreme conditions we simulate an intensive load by filling every block to its maximal size. In practice, in every
round, if a node does not have a full block to transmit, the node fills the block with random transactions, up to its maximal capacity, and then disseminates the full block.

In the current version, FireLedger uses a clique overlay for disseminating both blocks and headers. To avoid clogging the network, FireLedger has a basic flow control mechanism that prevents nodes from sending new blocks if the network is overloaded or if they have already disseminated enough blocks that have not been decided yet. Due to its modular design, a more sophisticated mechanism can be plugged into the system. This is left for future work.

7.2.1 FLO’s Throughput
To test FLO’s throughput we measured the blocks per second (bps) and the transactions per second (tps) rates. For each configuration we run the experiment for 3 minutes. The results were collected from all nodes and we took the average among them.

**BPS Rate** Due to the separation of blocks and headers, FLO’s throughput is mostly bounded by its bps rate. Figure 6 presents FLO’s bps rate for different $n$ and $\omega$ values. As expected, increasing $\omega$ yields higher bps due to better CPU utilization. In contrast, increasing $n$ decreases the bps because each decision requires more communication. Even so, FLO delivers thousands of bps under the majority of the tested configurations. Thus, for any configuration, FLO’s throughput is bounded by

$$tps \leq \beta \cdot bps.$$

**TPS Rate** We tested FLO’s throughput while varying $n, \omega, \sigma$ and $\beta$ values in the ranges described in Table 2. Figure 7 shows FLO’s throughput with the above configurations. Except few configurations where $\beta = 10$, FLO achieves between tens to hundreds of thousands of tps, depending on the specific configuration. Especially, with $\sigma = 512$, which according to [11] is the average size of a Bitcoin’s transaction, FLO peaks at around 160K tps even with $n = 10$. As expected, for larger $\sigma$ the performance decreases because less of the network’s bandwidth remains available for the headers, which limits the bps rate. It can be observed that the performance for large blocks with $n \in \{7, 10\}$ is better than when $n = 4$. This can be explained by the fact that the separation of blocks from headers allows for nodes in bigger clusters to collect more blocks that have not been decided yet. Thus, as the bps grows w.r.t. the number of workers, it respectively increases the tps rate. This does not manifest in small block sizes because with very small blocks there is little benefit from transmitting a block before its header. The experiment results demonstrate that FLO’s throughput is in line with the most demanding envisioned blockchain applications.
7.2.2 FLO’s Latency

To evaluate FLO’s latency, we measured the time it takes for a full block to be delivered by FLO. This time includes disseminating the block, its header, and to wait until it can be delivered by the round robin between FLO’s workers. We focus on the configurations in which $\sigma = 512$ (same as Bitcoin transactions). Figure 8 shows CDF charts for the various configurations. As expected, with $\omega = 1$ FLO’s latency is minimal and is less than 1 second even with $n = 10$ and $\beta = 1000$. Increasing $\omega$ results in an increase in the latency respectively. This is due to the fact that even a single worker’s delay is reflected in all, due to FLO’s round robin. Yet, even with $n = 10, \omega = 10$ and $\beta = 1000$ the latency in a single data-center deployment is below 8 seconds.

In order to understand the main bottlenecks of FLO, we divided every round of the algorithm into 5 different events: (A) block proposal, (B) header proposal, (C) tentative decision, (D) definite decision, and (E) delivering by FLO. Finally, we measured the time between each pair of consecutive events.

Figure 9 shows the relative execution time between each two consecutive events. It is easy to see that due to the separation of blocks and headers, the majority of time is spent between receiving a block and receiving its header. In addition, for $\omega > 1$, the workers cause an increase in the latency, as even a single worker’s delay, delays the whole system. Finally, increasing $n$ as well as $\beta$ causes the blocks dissemination event to take longer. This is despite using a clique layout. Other methods (e.g., gossip) may improve the throughput but not the latency.
Figure 8: CDF charts for $\sigma = 512$, $n \in \{4, 7, 10\}$, $\omega \in \{1, 5, 10\}$ and $\beta \in \{10, 100, 1000\}$ in a single data-center

7.3 FLO’s Scalability

To test FLO’s scalability we deployed a single data-center cluster of $n = 100$ machines and tested FLO’s tps rate with $\sigma = 512$, $\beta \in \{10, 100, 1000\}$ and $1 \leq \omega \leq 5$. We ran each configuration for 3 minutes. Figure [10] depicts the benchmark’s result. Thanks to FireLedger’s frugal communication pattern, as long as the fault free execution path take place, FLO can achieve around 60K tps (in a single data-center deployment). It can be seen that due to the cluster size, the number of workers has no effect because of the relatively large amount of communication that even a single worker consumes.

7.4 FLO Under Failures

7.4.1 Benign Failures

We tested FLO while suffering from crash failures of $f$ nodes (yet, we maintain $n = 3f + 1$ even in this benign case). Here, all faulty nodes crash in the middle of a run (such a node crashes with all of its workers), but the measurements are taken after the faulty nodes crash. Every benchmark was ran for 3 minutes and we calculated the average tps among the correct nodes. Figure [11] depicts FLO’s tps rate under various configurations while facing crash failures.

It can be seen that due to the full BBC phase that is now needed during faulty nodes rounds, for larger $n$ the tps is decreasing. Yet, FLO still reaches tens of thousands of tps despite these benign failures. This is due to the OBBC protocol and the basic failure detector described in Section 6.1.1.
7.4.2 Byzantine Failures

To test FLO’s performance when facing Byzantine failures, we deployed a Byzantine FLO node that operates as following: When started, every worker divides the cluster into two random parts, and for every given round it distributes different versions of the block to each part. Notice that in practice, invoking the recovery procedure may cause the nodes to become un-synchronized with each other, a fact that affects the performance as well. This increases the variance between measurements of the same settings. Hence, to be able to perform more measurements of each data point, for each configuration we run a series of short benchmarks (between 1 - 2 minutes each).

Figure 12 presents the tps rate for FLO when facing Byzantine failures w.r.t. the number of workers and the number of recoveries per second (rps). Smaller values of \( \beta \) and \( n \) imply more recovery events. Recall that during the recovery nodes halt. Thus, the above is expected due to the fact that each recovery ends faster when \( \beta \) and \( n \) are small. Yet, for bigger \( \beta \), the batching effect compensate for the small amount of recoveries and the long halts. The reason why for \( n = 10, \beta = 1000 \) and \( \omega = 5 \) the performance decreased so much is the underlying Byzantine consensus layer (BFT-SMaRt), which has to handle a large amount of data. To conclude, FLO delivers more than 10K tps in some scenarios even when facing Byzantine failures. Although the performance is lower than in optimistic executions, these type of failures are expected to be rare in permissioned blockchain clusters. And even with \( n = 10 \), if we set \( \beta = 1000 \) and \( \omega = 3 \) we obtain about 6K tps, roughly twice the average tps of VISA. Hence, FLO presents an attractive trade-off between scalability, performance and security.
We also tested FLO in a geo-distributed setting with nodes spread around the world. The nodes were placed, one node per region, by the following order, in Amazon’s Tokyo, Central, Frankfurt, Paris, Sao-Paulo, Oregon, Singapore, Sydney, Ireland and Ohio data-centers. We tested only fault free scenarios. Thus, we kept using BFT-SMaRt rather than its geo-distributed optimized version named WHEAT [70].

### 7.5 FLO in a Multi Data-Center Cluster

We tested FLO in a geo-distributed setting with nodes spread around the world. The nodes were placed, one node per region, by the following order, in Amazon’s Tokyo, Central, Frankfurt, Paris, Sao-Paulo, Oregon, Singapore, Sydney, Ireland and Ohio data-centers. We tested only fault free scenarios. Thus, we kept using BFT-SMaRt rather than its geo-distributed optimized version named WHEAT [70].

#### 7.5.1 FLO’s Throughput

As before, we first measured the bps rate for $n \in \{4, 7, 10\}$. Figure 13 depicts the bps rate for varying cluster sizes. As expected, due to lower network performance, the bps is less than 10% of its rate in single data-center clusters.

As in the single data-center deployment, we simulated high load by creating random transactions and run FLO with the following configurations: $n \in \{4, 7, 10\}, 1 \leq \omega \leq 10, \sigma = 512, \beta \in \{10, 100, 1000\}$. Every benchmark was run for 3 minutes. Figure 14 presents the benchmarks result. Obviously, tps is increasing with $w$ and $\beta$ due to a better CPU utilization and a batching effect. As for the increase of the tps with $n$, this is, as before, thanks to the separation of blocks from headers which allows bigger clusters to collect more blocks to decide on, thereby enhancing performance.

#### 7.5.2 FLO’s Latency

We tested FLO’s latency in the above multi data-center deployment. As before, we measured the time that takes a block to be delivered by FLO from the moment it was proposed by its creator. To avoid outliers,
Figure 12: FLO’s tps rate under Byzantine failures for $\sigma = 512, 1 \leq \omega \leq 5, \beta \in \{10, 100, 1000\}, n \in \{4, 7, 10\}$ and $f \in \{1, 2, 3\}$. The lines show the tps rate and the bars shows the recovery per second (rps) rate.

Figure 13: FLO’s bps rate for $n \in \{4, 7, 10\}$ in a multi data-center cluster.

7.6 FireLedger vs. Leading Alternatives

To the best of our knowledge, the current best performing alternative to FireLedger is HotStuff [74]. As we had no access to the codebase of HotStuff, we compare the declared performance of HotStuff from [74] with our own measurements of FireLedger using the exact same environment as in [74], namely c5.4xlarge AWS machines (16 vCPUs, 32 GiB RAM). We also compare with the numbers obtained for BFT-SMaRt [20], the previous state-of-the-art system, in this same setting. Figure 16 and 17 show this comparison’s results w.r.t. the $n$ and $\sigma$ parameters. For all cluster sizes FLO was deployed with $\beta = 1000$ and $\omega = 8$ with maximal resiliency $f = \lfloor n/3 \rfloor - 1$. In terms of throughput, for any $\sigma$ and $n$ FLO performs 20% – 300% better than HotStuff and 40% – 600% better than bft-SMaRt. Notice that HotStuff is implemented in C while FLO is in Java.

As for latency, due to the $f + 1$ finality of FireLedger and the fact that we run with maximal resiliency, FLO’s latency rises with to $n$. In contrast, transactions’ finality with both HotStuff and bft-SMaRt is at most three rounds (HotStuff), so their latency is much less impacted by $n$. Still, in all cases the latency obtained by FLO is better than most existing cryptocurrencies (and tokens), including the very recent Algo...
Figure 14: FLO’s transactions throughput for $\sigma = 512$ under various configurations in a multi data-center (Algorand) [49]. Libra’s target 10 seconds finality [43] is met by FLO when $n \leq 30$ nodes.

**Conclusions** The performance gap between FLO and the alternatives narrows as transactions become larger. This is because in such cases, the basic need to disseminate the transactions dominates the communication overhead, so a clever consensus protocol has less room for impact. Hence, one should consider compressing the data for large transactions. Also, by employing scalable Byzantine dissemination protocols for the transactions data, e.g., [39], FLO is likely to better handle large clusters.

8 Discussion

FireLedger is a communication frugal optimistic blockchain algorithm targeting environments where failures rarely occur. For example, FireLedger is likely to be very attractive for the FinTech industry, which uses highly secure and robust systems. FireLedger leverages blockchain’s iterativity as well as its cryptographic features to achieve its goal.

Loosely speaking, FLO employs FireLedger as a blockchain-based consensus algorithm rather than consensus-based blockchain. Our performance results show that it matches the requirements of real demanding commercial applications even when executing on common non-dedicated infrastructure.

In this paper we studied the distributed agreement aspects of FLO/FireLedger’s optimistic approach. In the future, we intend to examine other aspects such as sharding, data model, scalability, and improvement of FireLedger’s latency. Finally, our prototype implementation of this work is available in open source [10].
Figure 15: FLO’s latency in a multi data-center deployment for $\sigma = 512, \omega \in \{1, 5, 10\}, \beta \in \{10, 100, 1000\}$ and $n \in \{4, 7, 10\}$

Figure 16: Comparison of FLO and HotStuff on c5.4xlarge AWS machines and $f = \lfloor n/3 \rfloor - 1$
Figure 17: Comparison of FLO and bft-SMaRt on c5.4xlarge AWS machines and $f = \lfloor n/3 \rfloor - 1$
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A Optimistic Binary Byzantine Consensus

A.1 Introduction

The Optimistic Binary Byzantine Consensus (OBBC) is a key abstraction in WRB’s ability to deliver messages in a single communication step in favorable conditions, which is the basis for FireLedger’s low latency and frugal communication pattern (see lines 3–13 in Algorithm 1). Recall that the favorable conditions for FireLedger are that there is no Byzantine activity and the network delivers the proposer’s messages in a timely manner. If these conditions are met, all nodes receive the proposer’s message (line 7 in Algorithm 1). Consequently, they will all invoke OBBC with 1. In other words, the implementation of OBBC must terminate in a single round whenever all nodes vote 1. This enables us to develop a deterministic protocol that withstands $f < \frac{n}{3}$ byzantine nodes. This is in contrast to other existing OBBC implementations, e.g. [46, 62], that either rely on an oracle, randomization, or weaker failure model. To that end, we denote a BBC protocol that terminates quickly when all nodes vote for a given value $v$ by OBBC$_v$; specifically, Algorithm 1 invokes OBBC$_1$.

A.2 OBBC$_v$ Formal Definition

OBBC$_v$-Agreement and OBBC$_v$-Termination are the same as their BBC counterparts. In addition we define the following:

**evidence($v$)**: An evidence($v$) for a value $v$ is a cryptographic proof that can be verified by an external valid function. For a value $v' \neq v$, evidence($v$) = nil.

**OBBC$_v$-Validity**: A value $v'$ decided by a correct node was either proposed by a correct node or $v' = v$ and it was proposed by a node that has a valid evidence($v$).

**OBBC$_v$-Fast-Termination**: If no node has proposed $v' \neq v$, then $v$ will be decided by every correct node in a single communication step.

A.3 Implementing OBBC$_v$

The pseudocode implementation of OBBC$_v$ is listed in Algorithm 4. OBBC$_v$.propose receives two parameters: the actual proposal $v'$ and an evidence($v$). If $p$ is correct and $v = v'$, then evidence($v$) is valid. Else, if $v \neq v'$ then evidence($v$) is nil.

When $p$ invokes OBBC$_v$.propose($v'$, evidence($v$)), it performs the following:

- Broadcast $v'$ to all (line OH4).
- Waits for at most $n - f$ proposals. If a single value $v' = v$ has been received, then $p$ decides $v$ (lines OH5-OH8).
- Else, $p$ broadcast a request for evidence($v$) and waits for $n - f$ replies (which might be nil) (lines OH12-OH13 and lines OH23-OH24).
- In addition, if $p$ receives a request for evidence($v$) from $q$ and $p$ has a valid one, $p$ sends back that evidence to $q$ (lines OH20-OH21).
- Finally, if $p$ has received a valid evidence($v$) it adopts $v$. Then $p$ proposes its value through a regular BBC (lines OH15-OH19).
Algorithm 4: OBBC \( v \) - code for \( p \)

\begin{verbatim}
Procedure OBBC \( v \).propose\( (v',\\text{evidence}(v)) \)
assert \( v = v' \implies \text{evidence}(v) \text{ is valid} \);  
assert \( v \neq v' \implies \text{evidence}(v) = \text{nil} \);  
broadcast\( (v') \);  
wait until \( n - f \) proposals \( \hat{v} \) have been received;  
votes = \{\text{received proposals}\};  
if \( \text{votes} = \{v\} \) then 
\[ \text{decide } v; \]
\[ \text{return } v; \]
end  
evidences = \{\} /* couldn’t terminate quickly; run full protocol */  
broadcast\( (EV) \);  
wait until \( |\text{evidences}| = n - f \);  
/* see also line OB24 */  
new\( _v = v' \);  
if \( \text{evidences} \text{ contains a valid evidence}(v) \) then  
\[ \text{new}_v = v /* \text{notice: only } v \text{ can have a valid evidence} */ \]
end  
return \( \text{BBC}_v.\text{propose}(\text{new}_v) /* \text{invoke a BBC protocol} */ \)
upon receiving \( (EV) \) from \( q \) do  
\[ \text{send}(\text{evidence}(v)) \text{ to } q \]
end  
upon receiving an evidence\( (\hat{v}) \) from \( q \) do  
\[ \text{evidences} = \text{evidences} \cup \{\text{evidence}(\hat{v})\} \]
end  
if invoked Decide \( v \) and some node invoked \( \text{BBC}_v.\text{propose}(v') \) then  
\[ \text{BBC}_v.\text{propose}(v) /* \text{this is executed at most once} */ \]
end
\end{verbatim}

A.4 OBBC \( v \) Correctness proof

Lemma A.4.1. If a correct node \( p \) has decided \( v \) in line OB8 then all correct nodes who did not decide in line OB8 set \( \text{new}_v = v \).

Proof. By Algorithm 4, if \( p \) is correct and has decided at line OB8 then at least \( f + 1 \) correct nodes broadcast \( v \). Thus, at least \( f + 1 \) correct nodes have a valid \( \text{evidence}(v) \). As for \( v' \neq v \), \( \text{evidence}(v) = \text{nil} \) and by the ratio between \( n \) and \( f \), a correct node at line OB13 will receive at least one valid \( \text{evidence}(v) \) and thus set \( \text{new}_v = v \).

Lemma A.4.2. If a correct node \( p \) has decided \( v \) at line OB8 then all correct nodes will eventually decide \( v \).

Proof. By Lemma A.4.1 any correct node who did not decide at line OB8 sets \( \text{new}_v = v \). By BBC-Validity, if all correct nodes invoke \( \text{BBC}_v.\text{propose}(v) \) with the same value \( v \) then \( v \) must be decided.

Lemma A.4.3 (OBBC\( v \)-Agreement). No two correct nodes decide differently.

Proof. If any correct node has decided \( v \) at line OB8 then by Lemma A.4.2 all correct nodes will eventually decide \( v \). Else, all correct nodes decide at line OB19 and by BBC-Agreement no two correct nodes decide differently.

Lemma A.4.4 (OBBC\( v \)-Validity). A value \( v' \) decided by a correct node was either proposed by a correct node or \( v' = v \) and it was proposed by a node with a valid \( \text{evidence}(v) \).

Proof. Assume b.w.o.c. that \( v' \) has been decided by a correct node but had not been proposed by any correct node and no correct node has received a valid \( \text{evidence}(v') \). If no correct node proposed \( v' \), it means that all correct nodes executed lines OB13-OB19 and all invoked \( \text{BBC}_v.\text{propose} \) with their initial value. Further,
since none of them had a valid evidence($v'$), their initial value is $v' \neq v$, and therefore the same (there are only two possible values in BBC). Hence, by **BBC-Validity** all correct nodes decided $v'$ which is also the value proposed by all of them. A contradiction.

**Lemma A.4.5** (OBBC$_v$-Termination). Every correct node eventually decides.

**Proof.** By the ratio between $n$ and $f$, no correct node blocks forever at lines OB5 or OB13. By **BBC-Termination** no correct node is blocked forever at line OB19. Thus, every correct node eventually decides.

**Lemma A.4.6** (OBBC$_v$-Fast-Termination). If no process has proposed $v' \neq v$, then $v$ will be decided by every correct node in a single communication step.

**Proof.** By the ratio between $n$ and $f$ and by lines OB5–OB8 a correct nodes will receive $n - f$ proposals of $v$ and thus decide $v$ in a single communication step.

**Theorem A.4.1.** By Lemmas A.4.3, A.4.4, A.4.5 and A.4.6 Algorithm 4 solves the OBBC$_v$ problem.

A.5 OBBC$_v$ and WRB

In the context of WRB, we use OBBC$_v$.propose with evidence(1) = (m, sigproposer(m)) as an evidence to run the OBBC.propose procedure at lines 8–13 of Algorithm 1. As we expect to face mostly Byzantine failure free synchronized executions, it is beneficial to use OBBC$_1$. Even in the presence of benign failures, as long as the proposer is correct and messages from correct nodes arrive in a timely manner, OBBC$_1$ is still able to terminate fast in a single communication step.