

From Weakly-terminating Binary Agreement and Reliable Broadcast to Atomic Broadcast

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Abstract. We present a novel and simple solution to *Atomic Broadcast* (AB). We reduce AB to two subproblems. One of them is *Reliable Broadcast* (RB). We also introduce a subproblem we call *Weakly-terminating Binary Agreement* (WBA). WBA relaxes *Binary Agreement* (BA) protocols by not always terminating. WBA admits much simpler solutions than BA. We discuss concrete solutions to RB and WBA. We prove safety, liveness, and censorship resilience of our new AB protocol.

1 Introduction

The core task of an *Atomic Broadcast* (AB) system is to get a network to agree on a series of transactions. In the case of a database, these are user submitted database updates. In the case of blockchain, these are smart contract calls or transfers of digital tokens.

Different contexts lead to different solutions to Atomic Broadcast. The right solutions depend on how many nodes are allowed to be faulty and on the speed and reliability of message delivery. They also use different cryptographic primitives, providing different security models.

Many algorithms reduce Atomic Broadcast to some *broadcast* and *agreement* subprotocols. In these systems, Atomic Broadcast repeats the subprotocols over and over. First, one or more nodes broadcast values they propose for appending to the sequence next. Then, all nodes reach agreement on those values. After that, the nodes start over and broadcast again.

Atomic Broadcast protocols often need agreement subprotocols because ordinary broadcasting can be faulty. A proposing node may be malicious and send different values to different peers. Nodes can crash. Or the network could delay messages for very long periods of time.

Reliable Broadcast[3] (RB) protocols prevent malicious proposers and simplify agreement for Atomic Broadcast. In Reliable Broadcast, *correct* nodes will accept at most one proposal. But Reliable Broadcast protocols cannot help if a proposing node is offline. An agreement subprotocol needs to decide which proposed values to accept.

Binary Agreement[2] (BA) complements Reliable Broadcast by deciding which proposals to accept. In Binary Agreement, all correct nodes will output the same value of 0 or 1 at some point for a given proposal. If they all output 1, then the Atomic Broadcast protocol adds the proposal to the output sequence. If they all output 0 then the protocol drops the proposal.

Our contribution is this: we show we can replace Binary Agreement with a simpler problem. We call the new class of protocol *Weakly-terminating Binary Agreement* (WBA). In WBA, all correct nodes will output either 0 or 1, or they will not output at all. With WBA, if all nodes receive a correct proposal in time then they will output 1. If they all time out waiting for a correct proposal they output 0. Otherwise, unlike Binary Agreement, they may not output at all. This way Binary Agreement protocols are a subclass of Weakly-terminating Binary Agreement protocols.

Our Atomic Broadcast protocol is easier to implement because it uses WBA. This is because WBA protocols have fewer requirements than BA protocols and thus permits simpler solutions which we will present in this paper.

2 Network Model

We consider a distributed network consisting of *nodes* sending messages to each other. A known subset of n nodes are the *validators*. A node is *correct* if it executes the protocols as described, otherwise it is *faulty*. A faulty node can send any message and ignore the protocols, or even collude with other faulty nodes. We assume that strictly less than one third of the validators are faulty, i.e. our fault tolerance f satisfies $n > 3f$.

A *quorum* is a set of more than $q = \frac{n+f}{2}$ validators. Any two quorums intersect with more than f shared validators. So they always have a correct validator in common.

We assume that the network is *partially synchronous*:

All direct messages between correct nodes arrive eventually. There is a point in time called *Global Stabilization Time* (GST). After the GST messages arrive with maximum delay δ . We assume we know δ but we do not know GST. This assumption is for convenience. In [7] the case where we do not know δ is considered: By making all timeouts in the protocol increase over time, we can accommodate for the unknown delay.

3 Atomic Broadcast

Throughout the following sections we describe an Atomic Broadcast protocol. In an AB protocol some nodes are *proposers*. The proposers receive multiple inputs of some type \mathcal{V} . All nodes output multiple values of type \mathcal{V} so that:

- **Agreement:** If any correct node outputs v , every correct node will eventually output v .
- **Total Order:** If any correct node outputs v before w , all correct nodes output v before w .
- **Censorship Resilience:** Every input v to a correct proposer is eventually output by the correct nodes.

Together, Agreement and Total Order are summarized as *Safety*. The additional property that it does not stop outputting values is *Liveness*, which is implied by Censorship Resilience.

Instead of directly specifying which messages to send, we will solve the AB problem by using two sub-protocols and a timer. Our AB implementation will keep all inputs it received in a buffer, and proceed by making inputs to and processing outputs from the subprotocols.

4 Weakly-terminating Binary Agreement and Reliable Broadcast

In a *Weakly-terminating Binary Agreement* (WBA) protocol each validator receives at most one single-bit input (0 or 1) and makes at most one output and:

- **Agreement:** If one correct node outputs b , all correct nodes eventually output b .
- **Validity:** If the correct nodes output b , more than $q - f$ correct validators had input b .
- **Weak Termination:** If more than q correct validators have input b , the correct nodes eventually output b .

In *Reliable Broadcast* (RB) a designated *proposer* receives one input and each node makes at most one output and:

- **Agreement:** If one correct node outputs v , all correct nodes eventually output v .
- **Weak Termination:** If the proposer is correct and has input v , the correct nodes eventually output v .

WBA and RB have asynchronous solutions that do not rely on the partial synchrony assumption. But if there are upper bounds for message delays, these usually give rise to upper bounds for how much time WBA and RB take:

We say WBA *has delay* Δ if:

- If more than q correct validators get the same input before time $t \geq \text{GST}$, then all correct nodes output before $t + \Delta$.

- If any correct node outputs at time $t \geq \text{GST}$, then all correct nodes output before $t + \Delta$.

And RB has delay Δ if:

- If a correct proposer gets input before time $t \geq \text{GST}$, then all correct nodes output before $t + \Delta$.
- If any correct node outputs at time $t \geq \text{GST}$, then all correct nodes output before $t + \Delta$.

We present two algorithms for both RB and WBA in section 6.

5 Reducing AB to RB and WBA

5.1 Idea

Our Atomic Broadcast solution is leader-based, following [4,5]. The idea is to proceed in a sequence of rounds. Each round has a designated proposer node which is the leader. The leader proposes the next value for all nodes to output. The leader sequence could be pseudorandom or round-robin. What matters for liveness is that every proposer is the leader of infinitely many rounds.

We use an RB instance $\text{RB}[r]$ in every round r , so it is guaranteed that all nodes receive the same proposed value — or none, if the leader is faulty! To avoid waiting indefinitely for $\text{RB}[r]$ we use a timeout, and a WBA instance $\text{WBA}[r]$, where every validator inputs 1 if they receive an acceptable proposal in time or 0 if they hit the timeout while waiting. If $\text{WBA}[r]$ outputs 1, that is a decision to *finalize* it, i.e. return it as the next AB output.

Neither RB nor WBA guarantee that they output anything. But the crucial observation is that if $\text{RB}[r]$ does not output, all the correct validators will hit the timeout and input 0 in $\text{WBA}[r]$. In that case, $\text{WBA}[r]$ *does* output 0, because of Weak Termination. Thus the protocol can avoid getting stalled in round r by allowing the next leader in round $r + 1$ to make a proposal as soon as $\text{RB}[r]$ has output an acceptable proposal *or* $\text{WBA}[r]$ has output 0.

The devil is in the details: It can happen that $\text{RB}[r]$ does output an acceptable proposal and $\text{WBA}[r]$ outputs 0 anyway (because the timeout hit first for too many validators). But all nodes need to agree on whether the round- r proposal should be output before outputting the one from $r + 1$. That's why the proposal in round $r + 1$ needs to specify whether round r should be skipped.

5.2 The Algorithm

For each round $r \in \mathbb{N}$, there is an RB instance $\text{RB}[r]$ with the designated round- r leader as proposer, and a WBA instance $\text{WBA}[r]$. Values in $\text{RB}[r]$ are pairs (v, s) that roughly mean: “I propose outputting v right after the value from round s .” If s is omitted — we use the symbol \perp — that means it is proposed as the first value. Values in $\text{WBA}[r]$ are 0 or 1, where 0 means: “I did not get a proposal in time and vote to allow the next proposer to skip round r .” And 1 means: “I got a proposal and vote for finalizing it.”

We call round r *committed* if $\text{WBA}[r]$ has output 1, and *skippable* if $\text{WBA}[r]$ has output 0.

For any pair (v, s) with $s \in \mathbb{N} \cup \{\perp\}$, we call s the *parent round*; if $s \neq \perp$ and $\text{RB}[s]$ has an output, we call that the *parent* of (v, s) . The *ancestors* of (v, s) are its parent and all ancestors of its parent. A pair (v, \perp) has no parent or ancestors. If (v, s) with $s \neq \perp$ is the output of $\text{RB}[r]$, we also call s the *parent* of r . The *ancestors* of a round are its parent and all of its parent's ancestors.

\perp is *fertile* in round $r \in \mathbb{N}$ if all rounds $t < r$ are skippable. Some $s \in \mathbb{N}$ is *fertile* in round r if $s < r$, $\text{RB}[s]$ has an output with a fertile parent round and all rounds t with $s < t < r$ are skippable. If $\text{RB}[r]$ has an output with a fertile parent round, then that output is *accepted* in round r .

If (v, s) is accepted in round r and r is committed, then (v, s) and all its ancestors are *finalized*. We also call r and all of (v, s) 's ancestor rounds finalized. (Note that not all of those ancestor rounds are necessarily committed.)

The *current round* is the lowest round r which is neither skippable nor has an accepted value.

We assume that RB and WBA have delay Δ after GST.

We formulate the protocol by specifying what actions to take whenever certain conditions become true in a node N . These need to be checked whenever the timer fires or one of the subprotocols outputs. When we write “input” in some RB/WBA, we mean “input unless we have already made an input earlier”.

- When N is leader in r , and has an input v that has not been finalized yet, and there is a fertile s in r , it inputs (v, s) in $\text{RB}[r]$.
- When a new round r becomes current, N cancels and restarts the timer, with delay 2Δ .
- When the timer fires and r is current, N inputs 0 in $\text{WBA}[r]$.
- When there is an accepted value in a round r , N inputs 1 in $\text{WBA}[r]$.
- When new values become finalized, N outputs them, lower rounds first.

5.3 Pseudocode

For simplicity we assume that all subprotocols ignore unexpected inputs. Equivalently, $S.\text{input}(x)$ means we input x into subprotocol S , if it expects an input, i.e. if we have not made an input before and, in the case of RB, if we are the designated proposer. Outside the subprotocols, only three variables are needed. Upon startup, we initialize the timer and call the main handler:

```
def initialize():
    current:  $\mathbb{N}$  = 0
    undecided_round:  $\mathbb{N}$  = 0
    inputs: List[ $\mathcal{V}$ ] = []

    start_timer(2 *  $\Delta$ )
    on_subprotocol_output()
```

We first implement some definitions and a helper function for finalizing values in the right order:

```
def fertile( $r$ :  $\mathbb{N}$ ,  $s$ : Option[ $\mathbb{N}$ ])  $\rightarrow \mathbb{B}$ :
    if  $s == \perp$ :
        return  $\forall t < r. \text{WBA}[t].\text{output} == 0$ 
    else:
        return  $s < r \wedge (\forall u. s < u < r \rightarrow \text{WBA}[u].\text{output} == 0) \wedge \exists t. \text{accepted}(s, t)$ 
```

```
def accepted( $r$ :  $\mathbb{N}$ ,  $s$ : Option[ $\mathbb{N}$ ])  $\rightarrow \mathbb{B}$ :
    return fertile( $r, s$ )  $\wedge \exists v. \text{RB}[r].\text{output} == (v, s)$ 
```

```
def finalize( $r$ :  $\mathbb{N}$ ):
    ( $v, t$ ) =  $\text{RB}[r].\text{output}$ 
    if  $t \neq \perp \wedge t \geq \text{undecided\_round}$ :
        finalize( $t$ )
    inputs.remove( $v$ )
    output( $v$ )
```

Finally the main part of the protocol: a handler for inputs, i.e. user-submitted transactions, one for the timer, and one for whenever any WBA or RB subprotocol instance outputs a value.

```
def on_input( $v$ :  $\mathcal{V}$ ):
    inputs = [ $v$ ] + inputs
```

```
def on_timeout():
     $\text{WBA}[\text{current}].\text{input}(0)$ 
```

```

def on_subprotocol_output():
    if RB[current].output  $\neq \perp \vee$  WBA[current].output == 0:
        current = current + 1
        restart_timer(2 *  $\Delta$ )

    if len(inputs) > 0  $\wedge \exists s \in \text{Option}[\mathbb{N}]. \text{fertile}(\text{current}, s)$ :
        RB[current].input((inputs[0], s))

    if  $\exists s \in \mathbb{N}, t \in \text{Option}[\mathbb{N}]. \text{accepted}(s, t)$ :
        WBA[s].input(1)

    if  $\exists s \in \mathbb{N}, t \in \text{Option}[\mathbb{N}]. (s \geq \text{undecided\_round} \wedge \text{accepted}(s, t) \wedge \text{WBA}[s].\text{output} == 1)$ :
        finalize(s)
        undecided_round = s + 1

```

5.4 Proofs

We described the algorithm from the point of view of one node, or of someone who implements it. The proofs concern the behavior of a whole network, so we have to distinguish between the protocol states in different nodes at different times. We write r is (N, t) -committed to say node N at time t sees round r as committed. We write r is t -committed to say it is (N, t) -committed for every correct node N . Analogous definitions apply to the notions *skippable*, *accepted* and *finalized*.

Lemma 1. *Let N be a correct node, and let round r be (N, t) -committed.*

1. *Then r is (N, t') -committed for all $t' > t$.*
2. *There is a t' such that r is t' -committed.*
3. *If RB and WBA have delay Δ and $t \geq \text{GST}$ then r is $(t + \Delta)$ -committed.*

The same holds true for the properties skippable, accepted and finalized.

Proof. We only prove each statement for the committed case. A round r is (N, t) -committed when $\text{WBA}[r]$ has output 1. By definition, RB and WBA only output one item, showing (1). By the Agreement property of WBA, all correct nodes will eventually see the same output, which shows (2). Assuming we are after GST, and RB and WBA only take time Δ , we know that every correct node N' will be able to infer that same commitment within time Δ , as that happens through RB and WBA, proving (3).

The only aspect in which different correct nodes' views can differ is the order in which these properties are seen. However, finality was defined such that it does have a monotonicity property anyway, i.e. if a round r_2 becomes finalized later than r_1 , then $r_2 > r_1$:

Lemma 2. *Let N be a correct node.*

If both r_1 and r_2 are (N, t) -finalized they are either equal or ancestors of each other.

If r_1 is (N, t_1) -finalized and r_2 is (N, t_2) -finalized but not (N, t_1) -finalized then $t_2 > t_1$, $r_2 > r_1$ and r_1 is an ancestor of r_2 .

Proof. For the first claim we omit the prefix (N, t) since it's only about one node and one point in time.

That r_1 is finalized means that it is equal to or an ancestor of some round r'_1 that has accepted a value and is committed. Similarly r_2 is equal to or ancestor of a committed r'_2 with accepted value. If $r'_1 = r'_2$ then both r_1 and r_2 are ancestors of that round and therefore equal or ancestors of each other. So assume now that $r'_1 \neq r'_2$, w.l.o.g. $r'_2 > r'_1$. Let s be minimal among r'_2 and its ancestors such that $s > r'_1$. By the recursive definition of *accepted*, since r'_2 has an accepted value, so does s . Let (v, s') be accepted in s . If $s' > r'_1$ that

would contradict the minimality of s . If $s' < r'_1$ that would mean r'_1 must be skippable, which is impossible since r'_1 is committed. Therefore $r'_1 = s'$. Hence r'_1 is also an ancestor of r'_2 . Thus both r_1 and r_2 are ancestors of or equal to r'_2 , so they are ancestors of or equal to each other.

For the second claim note that $t_2 > t_1$ because the property of being finalized can only become true with more RB/WBA outputs arriving, not false again.

In particular both r_1 and r_2 are (N, t_2) -finalized. By the first part of the lemma that means they are ancestors of each other. If r_1 were greater than r_2 , r_1 by definition could not be (N, t_1) -finalized without r_2 also being (N, t_1) -finalized. Hence $r_2 > r_1$.

In other words the rounds are observed as finalized in increasing order; it cannot happen that a node first sees r as finalized, and then later $s < r$. It follows from the Agreement property of RB and WBA that the set of rounds that are eventually finalized is the same in all nodes. Hence in every correct node the k -th output is exactly the k -th element in the set of all rounds that eventually get finalized. That proves:

Theorem 1 (Safety; Agreement and Total Order). *If any correct node's k -th output is v , then every correct node will eventually output k values and the k -th one is v .* \square

So far we did not use the partial synchrony assumption, which indeed is not needed for safety. From now on assume that RB and WBA have delay Δ .

Lemma 3. *If all correct nodes start before GST, then for every r , all rounds $s < r$ are $(\text{GST} + 3r\Delta)$ -skippable or have a $(\text{GST} + 3r\Delta)$ -accepted value. In particular, at time $\text{GST} + 3r\Delta$, every correct node's current round is $\geq r$.*

Proof. We show this by induction on r .

The base case is trivial since there are no rounds before the first one.

So let $r > 0$ and $t = \text{GST} + 3r\Delta$, and assume the induction hypothesis: All rounds $s < r - 1$ are $(t - 3\Delta)$ -skippable or have a $(t - 3\Delta)$ -accepted value.

That means all correct validators start their timer for round $r - 1$ before $t - 3\Delta$.

If round $r - 1$ has a $(N, t - \Delta)$ -accepted value for some correct N , it has a t -accepted value.

Otherwise all correct validators will hit the timeout before $t - \Delta$ and input 0 in $\text{WBA}[r - 1]$, so WBA outputs 0 before t and $r - 1$ is t -skippable.

By Lemma 1, all rounds $s < r - 1$ are also t -skippable or have a t -accepted value, proving the statement for r .

Lemma 4. *Let R be the highest round that is current in any correct node at GST. Every round $r > R$ with a correct leader that has an unfinalized input eventually has an accepted value and becomes committed.*

Proof. By Lemma 3, all correct nodes eventually reach a round $\geq r$. Let N be the first correct node to do so, i.e. the node with the minimal t such that all rounds $s < r$ are (N, t) -skippable or have an (N, t) -accepted value. So no correct node starts its timer for round r before time t and no correct node inputs 0 in $\text{WBA}[r]$ before $t + 2\Delta$.

By our assumption $t > \text{GST}$, so by Lemma 1 all correct nodes reach round $\geq r$ before $t + \Delta$.

In particular if the leader L in round r is correct and has an input v that is not finalized yet, it will propose it before $t + \Delta$, and $\text{RB}[r]$ will output it in all correct nodes before $t + 2\Delta$. Since the parent of the proposal is $(L, t + \Delta)$ -accepted, it will be $(t + 2\Delta)$ -accepted. Thus the proposal itself is also $(t + 2\Delta)$ -accepted, and each correct validator inputs 1 in $\text{WBA}[r]$, so r becomes committed.

The requirement for the leader sequence was that every proposer is the leader infinitely many times, so Lemma 4 implies:

Theorem 2 (Liveness; Censorship Resilience). *Every value input in a correct proposer is eventually output by the correct nodes.* \square

6 Reliable Broadcast and Weakly-terminating Binary Agreement Solutions

6.1 Bracha's Algorithm

Bracha presents a simple implementation of Reliable Broadcast in [3]. His algorithm may be summarized as follows:

1. On input v , the proposer sends $(\text{initial}, v)$ to all nodes.
2. Each correct validator waits for either one $(\text{initial}, v)$ from the proposer, or a quorum of (echo, v) or (ready, v) from $> f$ validators, then sends (echo, v) to everyone.
3. Each correct validator waits for either a quorum of (echo, v) or (ready, v) from $> f$ validators, then sends (ready, v) to everyone.
4. Each correct node waits for (ready, v) from $> 2f$ validators, then outputs v .

A slight change turns this into a solution for the WBA problem as well:

1. Each correct validator waits for either an input b from the proposer, or a quorum of (vote, b) or (ready, b) from $> f$ validators, then sends (vote, b) to everyone.
2. Each correct validator waits for either a quorum of (vote, b) or (ready, b) from $> f$ validators, then sends (ready, b) to everyone.
3. Each correct node waits for (ready, b) from $> 2f$ validators, then outputs b .

The votes serve the same purpose as the echos, but instead of the value received from the proposer, they contain the sender's own input value.

These algorithms require a network where all correct validators are directly connected to all correct nodes, but work without any synchrony assumptions.

If the maximum message delay is δ , they have delay 3δ resp. 2δ .

6.2 Gossiping Quorums of Signatures

A different solution that works even in larger networks where a validator cannot expect to be directly connected to all other correct nodes is using a gossip mechanism to disseminate signatures and wait for a quorum. Gossip protocols are a broad subject themselves, so assume we have one that satisfies the following properties:

- If a correct node gossips a message m , eventually all correct nodes receive m .
- If any correct node receives a message m , eventually all correct nodes receive m .

Note that even if all correct nodes were connected to each other, this is a stronger guarantee than we get by just sending m to everyone: If a faulty node sends m to some and m' to other correct nodes, a gossip protocol guarantees that all correct nodes will receive both m and m' .

We also need cryptographic signatures, and a setup where each validator has a public key known to all nodes. In the following we write (\dots, σ) for a message where σ is a signature of the other fields.

Given these tools, a solution to RB is simply:

1. On input v , the proposer signs and gossips $(\text{initial}, v, \sigma)$.
2. When a correct validator receives $(\text{initial}, v, \sigma)$ signed by the proposer, and has not signed an echo yet, it signs and gossips $(\text{echo}, v, \sigma')$.
3. When a correct node receives $(\text{echo}, v, \sigma_i)$ with a set of signatures σ_i from a quorum of validators, it outputs v .

And WBA:

1. On input b , a correct validator signs and gossips (`vote`, b , σ).
2. When a correct node receives (`vote`, b , σ_i) with a set of signatures σ_i from a quorum of validators, it outputs b .

To see that these algorithms indeed solve the RB/WBA problems, observe that each correct validator only ever signs one `echo` or `vote` message, and since any two quorums overlap in at least one correct validator, there can be a quorum for at most one value. By our assumptions about the gossip protocol, if any correct node sees such a quorum, all of them will eventually see it. That implies RB and WBA Agreement. WBA Validity follows because a quorum must contain signatures from at least $q - f$ correct validators. RB and WBA Weak Termination follow because the gossip protocol guarantees to deliver a correct proposer's unique `initial` message to everyone, and a correct validator's `echo` or `vote`.

Note that the stronger guarantee that every correct node will receive all the messages received by at least one correct node, together with the use of cryptographic signatures, make one of the message rounds unnecessary compared to section 6.1. Thus the delay of WBA is just the time the gossip algorithm takes to deliver a message to all correct nodes, and the delay of RB is twice that.

By our assumption about gossip, if any correct node receives the `initial` message, all of them will. If the value v is large, the RB algorithm can thus easily avoid making lots of redundant copies by replacing the v in the `echo` messages with a cryptographic hash of v . Point 3 then has to be modified slightly, since a node can only output v once it has received a quorum of `echos` *and* the `initial` message.

7 Practical Considerations

7.1 Censorship Resilience with Unknown δ

The more realistic version of partial synchrony is with an unknown maximum message delay δ , which will result in unknown RB and WBA delays Δ . In theory all protocols can be adapted to that version by slowing down their clocks over time: the later a timer is started the more its delay is increased compared to the known- δ variant.

In practice, however, ever-increasing timeouts would mean that the user-visible delay caused by a single crashed or faulty proposer becomes longer and longer. So implementations usually reset the timeout back to a lower delay whenever a new value is finalized. If the timeout is reset to a very low value after each finalization, this breaks the proof of Theorem 2: It could be that not every (or even no) correct block proposer gets their proposals finalized, because the timer is always reset before it is their turn.

One strategy is to keep increasing the timeout until at least k out of p proposers got their most recent round committed. That achieves a lower level of censorship resilience: At least k proposers will eventually get all their inputs finalized. But if more than $p - k$ proposers are faulty, it will lead to the timeouts and user-visible delays increasing indefinitely.

Another approach is to modify the protocol: A round r is *orphaned* if it has accepted a value but there is a finalized round $s > r$ of which r is not an ancestor. So an orphaned round is one that we know will never get finalized. Now instead of tuples we use triples (v, r, s) as proposal values: The third entry specifies an orphaned *uncle* round, and if the proposal gets finalized, we output both the value accepted in s and the value v . (v, r, s) is only acceptable if s is \perp or has an accepted value. But in addition, we only input 1 in WBA if we either do not have an orphaned round or s points to the lowest orphaned round we know of. This effectively forces proposers in later rounds to acknowledge each orphaned value as an uncle and indirectly finalize it, too. Thus the algorithm remains fully censorship resilient even if the timeout is reset whenever a round is finalized. Note that this complicates the liveness proof, but since eventually all correct nodes will agree on which is the lowest orphaned round, it still works.

7.2 Validation

The validity of a proposed value often depends on its ancestors in practice: E.g. a client request to a database should only be executed once, so the proposal (v, r) is only valid if none of its ancestors already contains the value v . Or a smart contract call is only valid if the caller has enough funds to pay the transaction fees. Or a block in a blockchain must contain its parent's hash.

In order to account for these various details, one might extend the definition of accepted in order to only accept valid values. Be careful to avoid proposing values which would be considered invalid in your implementation, of course.

7.3 Spam

In theory the protocol runs an infinite number of RB and RV instances. But correct nodes make inputs to $\text{RB}[r]$ or $\text{WBA}[r]$ once a round $\geq r$ is current. Implementation should provide a way for correct nodes to reject incoming messages belonging to implausibly high rounds, e.g. by making the sender queue them until they are ready.

7.4 Minimum Delays

The protocol can be extended to respect a configured start time and a minimum delay between a value and its parent by adding two additional conditions for incrementing our round number and starting the timer: Round r becomes current when all rounds $s < r$ are skippable or have an accepted value *and*:

- The current time is at least the configured start time for the network.
- At least the minimum delay has passed since the highest timestamp of any accepted value.

8 Comparison with other Protocols

The protocol discussed in this paper has similarities to many other known protocols. HoneyBadgerBFT [10], DBFT [6] and Aleph [8] are examples of leaderless protocols that also use RB as a subprotocol.

HoneyBadgerBFT and DBFT also use full BA, which guarantees termination in all cases. As these protocols are leaderless, instead of one RB and BA instance per round, there is one instance per round *per proposer*. Assuming enough proposers are correct, it is guaranteed that a certain number of RB instances will terminate. This replaces the timer: Validators input 1 in a BA if the corresponding proposer's RB was among the first to deliver a value, otherwise they input 0. HoneyBadgerBFT then combines all proposed values where BA output 1, whereas DBFT just uses one of them. DBFT uses a BA solution that requires partial synchrony, whereas HoneyBadgerBFT is asynchronous thanks to a BA protocol based on threshold cryptography.

Aleph uses an asynchronous threshold cryptography-based BA in addition to RB, but reduces communication complexity by arranging the proposed values in a directed acyclic graph (DAG) where every proposal points to all proposals the sender has seen before, and works on an implicit message in several BA instances simultaneously.

While they share with our protocol the property of making subprotocols explicit, these three algorithms are vastly more complex than ours, which belongs to the family of leader-based partially synchronous protocols following the Lock-Commit paradigm [1]. It can be compared to a number of other members of that family, e.g. HotStuff [12], SBFT [9], Doomsbug [11], PBFT [5], Tendermint [4]. In this context, we will use the terms *block* and *value* synonymously.

SBFT trades a slightly lower fault tolerance for a fast-path that finalizes blocks in just one round of messages if *almost all* validators are correct. Doomsbug also finalizes blocks in a single round, but only with fault tolerance 1, and then later finalizes all blocks with a higher fault tolerance.

HotStuff and Libra support *pipelining*, similar to Aleph: To reduce overall communication complexity per block and to minimize the time between subsequent blocks, consensus for one of them takes several subsequent rounds to be reached. The downside is that the time to consensus is tied to how long it takes to broadcast the next proposals, as well as to any minimum time between blocks that one may want to enforce in practice.

PBFT and Tendermint do not do pipelining but try to finalize each block before the next one, which requires multiple rounds of messages.

By allowing the next proposal as soon as RB has output, but having a separate WBA instance per round, we do something in between: Very few RB messages have to be exchanged between subsequent blocks, but WBA can finalize the value without waiting for later rounds.

Apart from that, if the gossip-based RB and WBA implementations from section 6.2 are used, Tendermint is closest to our protocol. So we will compare them in detail:

First of all, the Tendermint paper describes a setting where a new instance of the protocol is run for every block, and the protocol outputs only once. But it can easily be adapted to make multiple outputs: Instead of re-proposing the valid value (see below), a proposer could create a new child of that value. For the purpose of our comparison we will work with that version.

In Tendermint, every round is subdivided into three phases:

- First the leader gossips a proposal, which also includes a value and the intended parent round — the proposer’s highest *valid round* (see below).
- Then all validators send a *prevote* with the hash of the proposal they received, or with *nil* if they received none.
- Finally the validators send a *precommit* with the hash of the proposal they saw a quorum of prevotes for, or *nil* if none.

Each phase has its own timeout: the first one is started when the round begins, but the prevote resp. precommit timeouts start when a quorum of prevote resp. precommit messages are seen, even if their content does not match.

The first two phases clearly correspond to RB, proposal to initial and prevote to echo. However there is no analog to a *nil* prevote in our protocol. Instead, RB simply does not guarantee to terminate, and there is one single timeout for it.

The third phase corresponds to WBA, precommit to vote, the proposal hash to 1 and *nil* to 0. The difference here is that Tendermint has a timer for this phase and the next proposer has to wait for another quorum or timeout even if there is a quorum of prevotes.

The *valid round* in Tendermint is the highest one in which we have seen a quorum of prevotes for the same proposal. It corresponds to the highest round with an accepted value. The *locked round* in a Tendermint node is the highest round for which *that node* has sent a precommit for a proposal. This is subjective and replaces the skipping rule: A node will refuse to prevote for a new proposal if our locked round is between it and its parent round.

9 Conclusion

We believe that isolating the RB and WBA subprotocols makes this partially synchronous Atomic Broadcast protocol particularly easy to understand, and gives intuitive meaning to its messages, without compromising on desirable properties regarding performance and fault tolerance.

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