Randomized Wait-Free Consensus using An Atomicity Assumption*

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Abstract. We present a randomized algorithm for asynchronous wait-free consensus using multi-writer multi-reader shared registers. This algorithm is based on earlier work by Chor, Israeli and Li (CIL) and is correct under the assumption that processes can perform a random choice and a write operation in one atomic step. The expected total work for our algorithm is shown to be $O(N \log(\log N))$, compared with $O(N^2)$ for the CIL algorithm and $O(N \log N)$ for the best weak-adversary algorithm previously known. We also model check instances of our algorithm using the probabilistic model checking tool PRISM.

Keywords: Asynchronous Consensus, Randomized Algorithms, Wait-Free Termination, Weak Adversary, Probabilistic Model Checking

AMS (2000): 68W15, 68W20, 68W40, 68Q25, 68Q60


1 Introduction

Distributed consensus refers to a class of problems in which a set of parallel processes exchange messages in order to agree on a common preference. Initially, each process is given an input value from a fixed, finite domain and, at the end of the algorithm, each non-faulty process outputs a decision value. Correctness requirements are typically formulated as follows.

- Validity: the output of any non-faulty process must have been the input of some process.
- Agreement: all non-faulty processes decide on the same value.
- Termination: every non-faulty process decides after a finite number of steps.

As shown in [FLP85], there exists no deterministic algorithm that solves distributed consensus in a setting of asynchronous communication with undetected process failure. Nonetheless, many efficient solutions exist under stronger assumptions (e.g. partial synchrony [DLS88] and failure detection [ACT00]) or weaker correctness requirements (e.g. probabilistic termination [CIL87]).


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Our algorithm falls into the category of randomized consensus algorithms\(^1\), where processes may use coin tosses to determine their course of actions. In this setting, termination is weakened to a probabilistic statement: the set of all non-terminating executions has probability 0. The first randomized consensus algorithm was proposed by Chor, Israeli and Li [CIL87,CIL94]. It satisfies the following termination condition.

- **Probabilistic wait-free termination**: with probability 1, each non-faulty process decides after a finite number of steps.

We adopt the same requirement. In fact, the logical structure of our algorithm closely resemble that in [CIL94], while we borrow ideas from [Cha96] to reduce the amount of shared and local data. We shall refer to [CIL94] as the original CIL algorithm and our own as the modified CIL algorithm.

**Adversary Models and Work Bounds.** To prove probabilistic termination, we must reason about probability distributions on the set of executions. This is done by specifying the so-called adversaries, which are fictitious entities designed to model scheduling uncertainties in a distributed environment. Mathematically, an adversary is a function mapping each finite execution to an available next operation. Such a function resolves all non-deterministic choices among parallel processes, thereby inducing a probability distribution on the set of executions. One can then ask if probabilistic termination is satisfied according to this distribution. By quantifying over all possible adversaries (of a certain strength), we obtain worst-case guarantees similar to those in a non-probabilistic setting.

The strength of an adversary varies according to the amount of information it can extract from a finite history. The strong adversaries have access to complete history of all processes and shared registers. Some weaker forms, such as write-oblivious and value-oblivious, delay the adversary’s knowledge of outcomes of internal coin tosses. The oblivious adversaries are the weakest, unable to observe any random outcomes and their subsequent effects on system dynamics.

Clearly, a stronger adversary model permits more possibilities and therefore renders consensus more difficult. Consensus against strong adversaries is shown to be \(\Omega(N^2/\log^2 N)\) in expected total work, where \(N\) is the number of processes participating in the algorithm [Asp98]. The best known algorithms achieve expected \(O(N^2 \log N)\) total work [BR91] and \(O(N \log^2 N)\) per process [AW96]. Against write-oblivious adversaries, one can achieve expected \(O(\log N)\) per process work and \(O(N \log N)\) total work [Aum97].

Our adversary model takes the form of an atomicity assumption\(^2\): processes can perform a random choice and a write operation in one atomic step. In particular, if the coin comes up heads, the process increments its round number and

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\(^1\) We refer to [Asp03] for a comprehensive overview on this topic.

\(^2\) Some authors refer to this assumption as local-oblivious, which we find misleading. Local-oblivious implies that local operations based on local coin tosses are also hidden from the adversary. This is stronger than we need, because in our case successful tosses are immediately written to shared memory, thus revealed to the adversary.
writes to the shared memory; otherwise, nothing happens. This model is weaker than the strong adversaries, but stronger than any other previously proposed adversary model.

The original CIL algorithm relies on the same atomicity assumption and achieves expected $O(N^2)$ total work [CIL94]. In the present paper, we replace the single-writer multiple-reader (SWMR) registers of [CIL94] with multi-writer multi-reader (MWMR) registers, thereby reducing the expected total work to $O(N \log(\log N))$.

Since our adversaries are value-sensitive, every non-faulty process must perform at least one read operation, otherwise we can easily construct an execution that violates the agreement property. Therefore, expected total work in this model is $\Omega(N)$, which almost matches our upper bound of $O(N \log(\log N))$.

We have adopted the worst case expected total work as our complexity measure, mainly because it is more natural to reason about the collective effect of all processes on the shared memory. In fact, per process work in our case is comparable to total work: if all but one process suffer crash failures, the lone survivor carries the total work burden and performs expected $O(N)$ tosses in order to pull far enough ahead for termination. In this sense, our algorithm is less efficient than [Cha96, Aum97], where polylogarithmic upper bounds are given for per process work. This is however a misleading comparison, because our adversary model is strictly stronger than those used in [Cha96, Aum97]. It is not known whether per process work is inevitably high in our particular setting.

**Probabilistic Model Checking.** We model check instances of our algorithm using PRISM, which can check PCTL (Probabilistic Computation Tree Logic) formulas against an MDP (Markov Decision Process) [PRI,BK98]. This tool has been applied to many randomized algorithms, including the consensus algorithm of Aspnes and Herlihy [AH90, KNS01] and Byzantine agreement [KN02].

Consensus algorithms are often hard to model check, because the state space grows exponentially with the number of participating processes. In [KNS01], PRISM is applied only to a shared-coin subroutine, while full correctness relies on verification using Cadence SMV, as well as higher level manual proofs. Unfortunately, the structure of our algorithm does not provide such convenient isolation of probabilistic reasoning. Nevertheless, we are able to build models of binary consensus with up to 4 processes and verify relevant properties.

In Section 6, we briefly describe these models and give a summary of PRISM results. In Section 7, we discuss our learning experience with PRISM and some prospects in improving feasibility of model checking.

**Overview.** Section 2 describes in greater detail our computational setting and assumptions. Section 3 presents the algorithm and correctness proofs are given in Sections 4 and 5. Section 6 is devoted to model checking and Section 7 contains closing discussions.
2 System Model

We consider a system of $N$ processes interacting asynchronously via shared memory objects. Each process $P_i$ is given as input an initial preference $p_0^i$, which belongs to a fixed, finite domain. Without loss of generality, this preference domain is assumed to be $\mathbb{Z}_K$ for some natural number constant $K \geq 2$. (As a convention, we write $\mathbb{Z}_K$ for $\{0, \ldots, K-1\}$ and $\mathbb{Z}_K^+$ for $\{1, \ldots, K-1\}$.)

We take a state-based view of our system. The local state of a process is determined by a valuation of all of its local variables, plus a program counter indicating the next line of code to be executed. In fact, it is trivial to include the program counter as an explicit variable, so that local state is fully determined by valuation of local variables. This is done in our PRISM models. The global state is then determined by local states of all $N$ processes, together with contents of shared memory objects. These objects are MWMR registers allowing primitive read and write operations. They are assumed to be linearizable [HW90], so that each memory operation can be viewed as taking place at a particular instant in time (as opposed to a time interval between invocation and response). Under this assumption, each read access returns the value written by the last write access in the linearized execution history.

A process executes a possibly infinite sequence of discrete steps, each consisting of a memory operation and/or a change in local state. It may also exhibit a limited form of non-deterministic behavior: crashing at any point of its execution. A crashed process may never recover and re-enter the algorithm.

An execution of the entire system is obtained by interleaving executions of individual processes, where scheduling among processes is determined by an adversary that satisfies the atomicity assumption stated in Section 1. That is, if a process is scheduled to toss a coin, it must be allowed to write to the memory (in case the coin lands heads) before another process is given a turn. The worst-case complexity is measured in terms of the expected number of read and write operations taken by all processes, quantifying over all admissible adversaries.

3 Modified CIL Algorithm

As in many other consensus algorithms [BO83,CIL94,AH90,Cha96], our processes advance through a sequence of rounds. During a particular round, a process goes through a possibly infinite sequence of phases, each of which is a complete pass through the main while-loop.

In original CIL, the shared memory is configured into an array of $N$ many SWMR registers, one for every process. Each register contains two pieces of information: round number $r_i$ and preference value $v_i$. At the beginning of each phase, process $P_i$ copies the contents of all register $j$ ($i \neq j$) and stores them locally. These entries are then examined to decide the next action of $P_i$: output a decision value and terminate, toss a coin to advance to the next round, jump to a higher round, etc.

This initial copying of each phase is the main source of inefficiency in original CIL: a large portion of copied data is inessential for decision making. For
example, \( P_i \) need not know exactly which \( P_j \) is in a higher round, as long as it knows some \( P_j \) is. This observation is precisely the motivation of our move from SWMR memory to MWMR memory. By treating processes anonymously, we reduce the number of read operations in the main loop from \( O(N) \) to \( O(1) \).

Following [Cha96], our MWMR shared memory is configured into \( K \) arrays of bits, each of length \( R + 2 \), where \( R := 2^{\lceil \log N \rceil} \). In other words, we have \( \text{mem} : \mathbb{Z}_{R+2} \times \mathbb{Z}_K \rightarrow \{0, 1\} \). (Recall that \( K \) is the size of the preference domain and is a constant, while \( N \) is the number of participating processes.) These bits can be interpreted as follows. For all \( r \in \mathbb{Z}_{R+1}^+ \) and \( v \in \mathbb{Z}_K \), \( \text{mem}(r, v) = 1 \) if and only if some process holds/held preference \( v \) while in round \( r \). These entries are thus initialized to 0. On the other hand, round-0 entries are initialized to 1 in order to avoid an initialization error\(^3\). Finally, round-(\( R + 1 \)) entries are initialized to 0 and are used for marking decision values. That is, if a process decides on value \( v \), it writes 1 to \( \text{mem}(R + 1, v) \).

As the algorithm progresses, a slow process (i.e., one with lower round number) always adopts the preference of one of the fastest processes it sees. Intuitively, we say the preference value of a slow process is eliminated by that of a fast process. (This notion is made precise in Definition 1 in Section 4.) Therefore the number of contending preference values never increases and the algorithm terminates when that number decreases to 1. If a process \( P_i \) sees itself at least two rounds ahead of any disagreeing peer, the algorithm guarantees that the preference \( p_i \) has eliminated every other contending value, therefore \( P_i \) can safely terminate with \( p_i \).

Biased coin tosses are used to break ties in the lead pack in such a way that, with probability 1, the number of contending preferences eventually reaches 1. This technique is used in [CIL94] and is quite different from the more common approach of shared coin subroutines, in which processes cast randomly generated votes to obtain a weak shared-coin [AH90,BR91].

Although every non-faulty process is guaranteed (with probability 1) to terminate after a finite number of steps, the round in which it terminates can become arbitrarily high. This requires an unbounded number of registers and is infeasible. Therefore we stop our algorithm when it reaches a certain round without successful termination, in which case we switch to a slower algorithm that requires bounded memory. We call this the exit algorithm. For convenience, the original CIL algorithm is chosen for this purpose\(^4\). We will show that any exit algorithm is invoked with probability at most \( \frac{1}{N} \), therefore the higher cost of original CIL does not affect overall expected complexity.

Figure 1(a) contains the pseudocode for process \( P_i \). The numbered lines can be described informally as follows.

(1) Check if some process has decided.

\(^3\) As noted in [CH05], original CIL contains this error: a process may decide for its own preference before a disagreeing process has a chance to write to \( \text{mem} \).

\(^4\) Technically, original CIL requires registers with unbounded size. However, according to [CIL94], the probability of non-termination is already extremely small \( (2^{-56}) \) when each register is 128 bits.
If so, decide for the same value.

3. Check if a disagreeing process has reached round \( r_i - 1 \).

4. If not, write 1 to \( \text{mem}(R + 1, p_i) \) and terminate with output \( p_i \).

5. Otherwise, if round \( R \) is reached, run the original CIL algorithm.

6. Otherwise, check if some process has reached round \( r_i + 1 \).

7. If not, advance to the next round with probability \( \frac{1}{2N} \).

8. Otherwise, run subroutine \text{Jump} to catch up with faster processes.

Notice that the atomicity assumption discussed in Section 1 applies to Line (7). This prevents the adversary from selectively delaying write operations of processes who are already to advance to the next round.

Figures 1(b) and 1(c) contain the subroutines \text{ReadMem} and \text{Jump}, respectively. The former is used to read from the shared memory, while the latter is used by a lagging process to catch up with the fastest processes. When called with parameters \( r \) and \( p \), \text{ReadMem} scans one-by-one the \( r \)-th entry of every bit vector, except for the \( p \)-th. In other words, we would like to know if any process has reached round \( r \) with preference other than \( p \). It returns the first \( k \) such that both \( k \neq p \) and, at the time of read access, \( \text{mem}(r, k) = 1 \). If no such \( k \) is encountered, \text{ReadMem} returns \( K \).

In every pass through the \textbf{while}-loop of Figure 1(a), \text{ReadMem} is called with at most three round numbers: \( R + 1 \), \( r_i - 1 \), and \( r_i + 1 \). This does not reveal the highest round ever reached by any process. Therefore, a separate subroutine \text{Jump} is run when the process sees itself behind. This is a key difference between our algorithm and original CIL: in exchange for fewer read operations in the main loop, more work is needed for slower processes to catch up.

The subroutine \text{Jump} can be implemented in various ways. A simple solution is to increment \( r_i \) by 1 and then call \text{ReadMem} once to find the least \( k \) such that \( \text{mem}(r_i, k) = 1 \). This requires a constant amount of work per invocation of \text{Jump} and is implemented in our PRISM models. However, a process lagging way behind may need to step through the main loop as many as \( R \) times in order to catch up. Hence we opt for a faster implementation, which is essentially a binary search on \text{mem}. This involves \( O(\log(\log N)) \) operations per invocation of \text{Jump}, but a lagging process can catch up with the leaders in one complete phase (provided no further progress is made in the mean time).

4 Validity and Agreement

In this section, we treat all coin tosses as non-deterministic choices. Let \( s_0 \) denote the initial state of our system, where all \( N \) processes as well as the shared memory have been properly initialized. A path of the system is a finite sequence of states \( s_0, s_1, \ldots, s_m \) where, for all \( j \in \mathbb{Z}_m \), \( s_{j+1} \) can be obtained from \( s_j \) by allowing exactly one non-faulty process to execute its next instruction. A state \( s \) is reachable if there is a path ending in \( s \). Finally, a value \( k \in \mathbb{Z}_K \) is said to be valid if there is \( i \in \mathbb{Z}_N \) such that \( k \) equals the input \( p_i^0 \) to process \( P_i \).

We use record notation to indicate valuation of variables. For example, \( s.r_i \) denotes the round number of \( P_i \) in state \( s \). If \( P_i \) is running a subroutine (e.g.
ModifiedCIL(i, p₀)
local variables
// round number
rᵢ ∈ ℤₘ₊₂,
// preference
pᵢ ∈ ℤₖ,
// decision value
dᵢ ∈ ℤₖ₊₁,
// values read from memory
aheadᵢ, behindᵢ ∈ ℤₖ₊₁
begin
    pᵢ := p₀ᵢ; rᵢ := 0;
    while rᵢ ≤ R do
        (1) dᵢ := ReadMem(R + 1, K);
        (2) if dᵢ ≠ K then return dᵢ;
            if rᵢ > 0 then {
                (3) behindᵢ := ReadMem(rᵢ − 1, pᵢ);
                (4) if behindᵢ = K then {
                    mem(R + 1, pᵢ) := 1;
                    return pᵢ;
                }
            }
            elseif rᵢ = R then return OriginalCIL(i, pᵢ)
        }
        (6) aheadᵢ := ReadMem(rᵢ + 1, K);
            if aheadᵢ = K then
                (7) with probability 1/2ᵐ do {
                    rᵢ := rᵢ + 1;
                    mem(rᵢ, pᵢ) := 1
                }
        (8) else {rᵢ, pᵢ} := Jump(rᵢ + 1, aheadᵢ)
    od
end

(a) Main Algorithm.

ReadMem(r, p)
local variables
// counter
k ∈ ℤₖ,
// preference value found
v ∈ ℤₖ₊₁,
begin
    k := 0; v := K;
    while k < K and v = K do
        if mem(r, k) = 1 and k ≠ p then
            v := k;
            k := k + 1
        od
        return v
end

(b) Subroutine ReadMem.

Jump(r, p)
local variables
// confirmed round and preference
r’ ∈ ℤₘ₊₁, p’ ∈ ℤₖ,
// current round and preference
l ∈ ℤₘ₊₁, u ∈ ℤₖ₊₁,
// counter
C ∈ ℤₘ₊₁,
begin
    if r ≥ R then return (r, p);
        r’ := r; p’ := p; c := [log(R − r)];
    while c > 0 do
        l := r’ + 2ᶜ⁻¹;
        if l ≤ R then {
            u := ReadMem(l, K);
            if u ≠ K then {
                r’ := l; p’ := u
            }
        }
        c := c − 1
    od
    return (r’, p’)
end

(c) Subroutine Jump.

Fig. 1. Modified CIL Algorithm
ReadMem), we add subscript $i$ to variables of that subroutine (e.g. $s.k_i$ and $s.v_i$). This should not cause any confusion because each process runs at most one instance of any subroutine at any given point of the execution.

First we state some properties about mem and subroutines ReadMem and Jump. Lemma 1 says that an entry in mem never changes from 1 to 0. Lemma 2 says that the return value of ReadMem is correct (although it may be out-of-date). Similarly, Lemma 3 states the correctness of Jump.

**Lemma 1.** Let $r \in \mathbb{Z}_{R+2}$, $v \in \mathbb{Z}_K$ and a path $s_0 \ldots s_m$ be given. Suppose there is $j \in \mathbb{Z}_{m+1}$ with $s_j.mem(r, v) = 1$. Then $s_j'.mem(r, v) = 1$ for all $j \leq j' \leq m$.

**Proof.** A process writes to the shared memory only if it executes Lines (4) or (7) in Figure 1(a). In either case, the value 1 is written. Therefore, once the return value of ReadMem is correct (although it may be out-of-date). Similarly, Lemma 3 states the correctness of Jump.

**Lemma 2.** Let $r \in \mathbb{Z}_{R+2}$, $p, v \in \mathbb{Z}_{K+1}$ and a path $s_0 \ldots s_m$ be given. If the last step is $\text{ReadMem}(r, p)$ returning $v \neq K$, then $s_m.mem(r, v) = 1$.

**Proof.** The return value $v$ of ReadMem is set to a value other than $K$ only if the if-then clause is executed. Let $s_j$ ($0 \leq j \leq m$) be the state from which this instance of ReadMem reads from mem$(r, v)$. Clearly, $s_j.mem(r, v) = 1$. By Lemma 1, this also holds in $s_m$.

**Lemma 3.** Let $r, r'' \in \mathbb{Z}_{R+1}$, $p, p'' \in \mathbb{Z}_K$ and a path $s_0 \ldots s_m$ be given. Suppose the last step is $\text{Jump}(r, p)$ returning $(r'', p'')$. If $\text{mem}(r, p) = 1$ when $\text{Jump}(r, p)$ is called, then $s_m.mem(r'', p'') = 1$.

**Proof.** We prove that $\text{mem}(r', p') = 1$ is an invariant of the while-loop in Jump. By assumption, the claim holds for initial values $r' = r$ and $p' = p$. Noticed that $\langle r', p' \rangle$ is updated only if the if-then clause is executed, in which case $v \neq K$. Since $(r', p')$ is the return value of ReadMem$(l, K)$, we have by Lemma 2 that $\text{mem}(l, v) = 1$, hence $\text{mem}(r', p') = 1$ still holds after the update. Let $s_j$ be the state immediately after the last update of $\langle r', p' \rangle$. Then we know $s_j.mem(r'', p'') = 1$. By Lemma 1, this also holds in $s_m$.

**Lemma 4.** Let a path $s_0 \ldots s_m$ be given.

(i) For all $i \in \mathbb{Z}_N$, $s_m.r_i \leq R$ implies $s_m.mem(s_m.r_i, s_m.p_i) = 1$.

(ii) For all $r \in \mathbb{Z}_{R+2}$ and $v \in \mathbb{Z}_K$, $s_m.mem(r, v) = 1$ implies there exist $i \in \mathbb{Z}_N$ and $j \in \mathbb{Z}_{m+1}$ such that $s_j.p_i = v$.

**Proof.** We proceed by induction on the length of paths. For the initial state $s_0$, recall that round-0 entries are initialized to 1 and all other entries 0, therefore the two claims hold trivially.

Now we consider a path $s_0 \ldots s_m s_{m+1}$. Suppose the last step is taken by process $P_i$. Let $r$ denote $s_m.r_i$ and $v$ denote $s_m.p_i$. Notice that only Lines (4), (7) and (8) in Figure 1(a) update variables $r_i, p_i$ and mem.
Line (4). By Lemma 1, Item (i) is trivial because $r_i$ is not updated. Item (ii) holds because the only entry of interest is $\text{mem}(R + 1, v)$ and by definition $v = s_{m+1}.p_i$.

Line (7). We may focus on the case in which the coin toss is successful. Recall that the two updates in Line (7) are assumed to be atomic. If $r + 1 < R$, then $r < R$ and $\text{mem}(r + 1, v)$ is set to 1. This proves Item (i). On the other hand, $s_{m+1}.p_i = v$, therefore Item (ii) also holds.

Line (8). Item (i) follows from Lemma 3. Item (ii) follows from the induction hypothesis.

\begin{theorem}
The following claims hold in every reachable state $s$.
\begin{enumerate}
  \item For every $i \in \mathbb{Z}_N$, $s.p_i$ is valid.
  \item For every $r \in \mathbb{Z}_{R+2}$ and $v \in \mathbb{Z}_K$, $s.\text{mem}(r, v) = 1$ implies $v$ is valid.
  \item For every $i \in \mathbb{Z}_N$, if $s.d_i \neq K$ then $s.d_i$ is valid. Similarly for $s.\text{ahead}_i$ and $s.\text{behind}_i$.
\end{enumerate}
\end{theorem}

\begin{proof}
We prove these claims simultaneously using induction on the length of paths. First consider the initial state $s_0$. For Item (i), every $s_0.p_i$ is valid because it is set to the input value $p_0^i$. Item (ii) holds trivially because all but round-0 entries are initialized to 0. Item (iii) is also trivial because all relevant variables are initialized to $K$.

Now we consider a path $s_0 \ldots s_{m}s_{m+1}$. Suppose the last is taken by process $P_i$. We examine Figure 1(a) line by line for all possible actions of $P_i$.

Line (1). In this case, one update is possible: $d_i$ is set to the return value $v$ of $\text{ReadMem}(R + 1, K)$. Suppose $v$ is not $K$. Then we can apply Lemma 2 to conclude that $s_{m+1}.\text{mem}(R + 1, v) = 1$. Since $\text{mem}$ is not updated in the last step, this also holds in $s_m$. Applying the induction hypothesis, we conclude that $s_{m+1}.d_i = v$ is valid.

Line (2). In this case, $P_i$ terminates by returning the value $s_m.d_i$ and there are no variable updates. We simply apply the induction hypothesis.

Lines (3) and (6). Similar to Line (2).

Line (4). In this case, $\text{mem}(R + 1, s_m.p_i)$ is set to 1. By the induction hypothesis, $s_m.p_i$ is valid. Therefore Items (ii) hold in $s_{m+1}$. Item (iii) follows from the inductive hypothesis.

Line (7). Similar to Line (4).

Line (8). $(r_i, p_i)$ is set to the return values of $\text{Jump}$. Notice that this update is executed only if $s_m.\text{ahead}_i \neq K$. Therefore, we can conclude that in Line (6) $\text{ReadMem}$ returned a value $v$ other than $K$. Moreover, notice that from then on $r_i$ and $\text{ahead}_i$ have not been updated. Therefore, by Lemmas 1 and 2, we know that $\text{mem}(s_m.r_i + 1, s_m.\text{ahead}_i) = 1$ at the time $\text{Jump}$ is called. Applying Lemma 3, we have $s_{m+1}.\text{mem}(s_{m+1}.r_i, s_{m+1}.p_i) = 1$. Since $\text{mem}$ is not updated in the last step, we have $s_{m}.\text{mem}(s_{m+1}.r_i, s_{m+1}.p_i) = 1$. Applying the induction hypothesis, we conclude that $s_{m+1}.p_i$ is valid.
\end{proof}
Corollary 1. The modified CIL algorithm in Figure 1 is valid, assuming the exit algorithm (in this case, the original CIL algorithm) is valid.

Next we prove agreement. A key ingredient is a predicate \( \Phi \) on global states.

Definition 1. Let \( v, v' \in \mathbb{Z}_K \) and \( r \in \mathbb{Z}_{R+1}^+ \) be given. We say that \( v \) eliminates \( v' \) in round \( r \) in global state \( s \) (denoted \( s \models \Phi(v, v', r) \)) just in case \( s.\text{mem}(r, v) = 1 \) and \( s.\text{mem}(r - 1, v') = 0 \).

We state a string of lemmas leading to the claim that no two processes terminating by Line (4) do so with conflicting decision values (Lemma 8). First, if an entry \( \text{mem}(r, v) \) is marked 1, then every entry \( \text{mem}(r', v) \) with \( r' \leq r \) is also marked 1 (Lemma 5). Second, if \( v' \) is eliminated by \( v \) in round \( r \), then no process subsequently reaches round \( r \) with preference \( v' \) (Lemma 6). Finally, if a process \( P_i \) terminates by Line (4) with value \( v \) in round \( r \), then every other \( v' \) must have been eliminated by \( v \) in round \( r \) at some earlier state (Lemma 7).

Lemma 5. Let \( s \) be a reachable state. For all \( r \in \mathbb{Z}_{R+1}^+ \) and \( v \in \mathbb{Z}_K \), if \( s.\text{mem}(r, v) = 1 \) then \( s.\text{mem}(r', v) = 1 \) for all \( r' \leq r \).

Proof. We proceed by induction on the length of paths. Clearly this holds at the initial state \( s_0 \). Consider a path of the form \( s_0 \ldots s_m \) and assume the property holds for \( s_m \). The only case of interest is when \( \text{mem}(r, v) \) changes from 0 to 1 as the result of some process \( P_i \) executes Line (7) from \( s_m \). In that case, we have \( s_m.r_i = r - 1 \) and \( s_m.p_i = v \). By Lemma 4, we may infer that \( s_{m+1}.\text{mem}(r - 1, v) = 1 \). By the induction hypothesis, \( s_{m+1}.\text{mem}(r', v) = 1 \) for all \( r' \leq r - 1 \). Using Lemma 1, we conclude that \( s_{m+1}.\text{mem}(r', v) = 1 \) for all \( r' \leq r \).

Lemma 6. Let \( v, v' \in \mathbb{Z}_K \) and \( r \in \mathbb{Z}_{R+1}^+ \) be given. Consider a path \( s_0 \ldots s_m \) such that \( s_j \models \Phi(v, v', r) \) for some \( j \in \mathbb{Z}_{m+1} \). Then, for all \( j' \in \{j, \ldots, m\} \), \( s_{j'}.\text{mem}(r, v') = 0 \).

Proof. By the definition of \( \Phi \), we have \( s_j.\text{mem}(r - 1, v') = 0 \) and \( s_j.\text{mem}(r, v') = 1 \). Using Lemma 5, we infer that \( s_{j'}.\text{mem}(r, v') = 0 \). For contradiction, suppose that the claim doesn’t hold. We focus on the least \( j' \geq j \) with \( s_{j'}.\text{mem}(r, v') = 1 \). Then it must be the case that \( s_{j'-1}.\text{mem}(r, v') = 0 \) and some process \( P_i \) executes Line (7) from \( s_{j'-1} \). Moreover, \( s_{j'-1}.r_i = r - 1 \) and \( s_{j'-1}.p_i = v' \).

On the other hand, using Lemma 4 and the fact that \( s_{j'.}\text{mem}(r - 1, v') = 0 \), we know either \( s_{j'.}r_i < r - 1 \) or \( s_{j'.}p_i \neq v' \). Therefore, \( P_i \) must have entered the current phase after \( s_j \). Since \( \text{mem}(r, v) \) is 1 in every state following \( s_j \), the invocation of \text{ReadMem} in Line (6) of the current phase of \( P_i \) must have returned a value other than \( K \). This contradicts the claim that \( P_i \) executes Line (7) in the current phase.

Lemma 7. Consider a path \( s_0 \ldots s_{m+1} \). Suppose that in the last step some process \( P_i \) terminates by executing Line (4). Let \( r \) denote \( s_m.r_i \) and \( v \) denote \( s_m.p_i \). For every \( v' \neq v \), there is \( j' \in \mathbb{Z}_{m+1} \) such that \( s_{j'} \models \Phi(v, v', r) \).
Proof. Let \( v' \neq v \) be given and let \( s_j \) denote the first state in which \( P_i \) enters the current phase. Thus \( s_j.r_i = r \) and \( s_j.p_i = v \). By Lemma 4 and Lemma 1, we have \( s_{j''}.\text{mem}(r,v) = 1 \) for all \( j'' \in \{j, \ldots, m\} \).

Since Line (4) is executed, \( r \) must be non-zero and the invocation of \text{ReadMem} \ in Line (3) must have returned \( K \). Let \( s_{j'} \) be the state from which \text{ReadMem} \ reads \( \text{mem}(r-1,v') \). Since the return value of \text{ReadMem} \ is \( K \), we may infer that \( s_{j'}.\text{mem}(r-1,v') = 0 \). Moreover, we have \( j' > j \) and hence \( s_{j'}.\text{mem}(r,v) = 1 \). Therefore \( s_{j'} \models \Phi(v,v',r) \).

Lemma 8. Let a path \( s_0 \ldots s_m \) and \( j, j' \in \mathbb{Z}_{m+1} \) be given. Assume that process \( P_i \) terminates by Line (4) with output \( v \) from state \( s_j \) and some other process \( P_{v'} \) does the same with output \( v' \) from state \( s_{j'} \). Then \( v = v' \).

Proof. For the sake of contradiction, suppose \( v \neq v' \). Let \( r \) and \( r' \) denote the final round numbers of \( P_i \) and \( P_{v'} \), respectively. Without loss of generality, assume that \( r \leq r' \). By Lemma 7, we know that \( s_j \models \Phi(v,v',r) \), therefore \( s_j.\text{mem}(r,v) = 1 \) and \( s_j.\text{mem}(r-1,v') = 0 \). On the other hand, \( s_{j'}.r_i = r' \) and \( s_{j'}.p_i = v' \), so by Lemma 4 we have \( s_{j'}.\text{mem}(r',v') = 1 \).

First we consider the case in which \( j < j' \). By Lemma 6, we know that \( s_{j}.\text{mem}(r',v') = 0 \). Applying Lemma 5, we have \( s_{j}.\text{mem}(r',v') = 0 \), which yields a contradiction.

Next we consider the case in which \( j' < j \). By Lemma 1, we may infer that \( s_{j}.\text{mem}(r',v') = 1 \). By Lemma 5, this implies \( s_{j}.\text{mem}(r-1,v') = 1 \), contradicting the fact that \( s_j \models \Phi(v,v',r) \).

It remains to consider termination by Line (2). Lemma 9 below implies that every process terminating by Line (2) must be preceded by a process terminating by Line (4) with the same decision.

Lemma 9. Let \( v \in \mathbb{Z}_K \) and a path \( s_0 \ldots s_m \) be given. Assume that \( s_m.\text{mem}(R+1,v) = 1 \). There is \( j \in \mathbb{Z}_{m+1} \) such that some process \( P_i \) terminates with decision value \( v \) by executing Line (4) from \( s_j \).

Proof. Let \( j \) be the index for the first state in this path such that \( s_j.\text{mem}(R+1,v) = 1 \). Since \( \text{mem}(R+1,v) \) is initialized to 0, we know that \( j > 0 \). Let \( P_i \) be the process writing to \( \text{mem}(R+1,v) \) from \( s_{j-1} \). Then \( P_i \) must have terminated with decision value \( v \) by executing Line (4) from \( s_{j-1} \).

Theorem 2. Let a path \( s_0 \ldots s_m \) be given. Assume that process \( P_i \) terminates by executing either Line (2) or Line (4) from state \( s_j \) \((j \in \mathbb{Z}_{m+1})\) and its decision value is \( v \). Similarly for \( P_{v'} \), \( s_j' \) and \( v' \). Then \( v = v' \).

Proof. We claim that there exist \( j'' \in \mathbb{Z}_{m+1} \) and \( i'' \in \mathbb{Z}_N \) such that \( P_{i''} \) terminates with decision value \( v \) by executing Line (4) from \( s_{j''} \). If \( P_i \) terminates by Line (4), then we simply set \( i'' := i \) and \( j'' := j \). Otherwise, \( P_i \) terminates by Line (2) and the invocation of \text{ReadMem} \ in Line (1) of the last phase of \( P_i \) must have returned \( v \neq K \). By Lemma 2 and Lemma 1, we know that \( s_m.\text{mem}(R+1,v) = 1 \). We can then choose \( j'' \) and \( i'' \) using Lemma 9.

The same claim also holds for \( v' \). Now we apply Lemma 8 to infer that \( v = v' \).
5 Probabilistic Termination and Expected Complexity

Let us first consider the amount of work required during each phase of the algorithm. (Recall that a phase is an entire pass through the while-loop in Figure 1(a)). Notice each phase involves at most (i) three calls to ReadMem, (ii) one write operation and (iii) one call to Jump. Each call to ReadMem requires $O(1)$ read operations, because the size $K$ of the preference domain is a constant in our analysis. Therefore, aside from Jump, each phase involves constant work.

Consider the while-loop in Jump. Each pass through this loop involves at most one call to ReadMem. Furthermore, this loop is executed at most $\log R + 1$ times. Since $R = 2[\log N]$ by definition, each call to Jump requires $O(\log(\log N))$ read operations. This is then also the cost of a complete phase. Later on, we will prove that the expected number of complete phases until at least one process terminates successfully is $O(N)$ and hence the expected number of read/write operations is $O(N \log(\log N))$ (Lemma 13).

For any state $s$, let $s.r_{\max}$ denote the highest round reached by any process in state $s$. In other words, $s.r_{\max} := \max_{i \in \mathbb{Z}_N} s.r_i$. Since the two updates in Line (7) of Figure 1(a) are performed in a single step, $s.r_{\max}$ is also the largest $r$ such that $s.mem(r, v) = 1$ for some value $v \in \{0, \ldots, K - 1\}$. Lemma 10 below says, if no leader advances to round $r_{\max} + 1$, a lagging process can catch up to round $r_{\max}$ in one complete phase. Lemma 11 then shows, whenever $r_{\max}$ is at most $R - 2$, the probability of at least one process terminating successfully within the next two rounds is bounded below by a constant. Moreover, this termination takes place before $15N$ complete phases are executed. The proof of Lemma 11 strongly resembles the analysis given in [CIL94].

Lemma 10. Let $s_0 \ldots s_m \ldots s_m'$ be a path with $m < m'$. Assume that $s_j.r_{\max} = s_m.r_{\max}$ for every $j \in \{m, \ldots, m'\}$. Moreover, assume that $P_i$ completes a phase between $s_m$ and $s_m'$ without crashing, successfully terminating or switching to the exit algorithm. Then $s_m'.r_i = s_m.r_{\max}$.

Proof. For readability, write $r_{\max}$ for $s_m.r_{\max}$ and $r$ for $s_m.r_i + 1$. Consider the first complete phase executed by $P_i$ between $s_m$ and $s_m'$. Without loss of generality, assume that $s_m$ is the first state in that phase and that $r \leq r_{\max}$.

By assumption, $P_i$ does not crash, terminate, or exit. Therefore it reaches Line (6) in this phase. By Lemma 1 and Lemma 5, $r \leq r_{\max}$ implies there is $v \in \mathbb{Z}_K$ such that $s_j.mem(r, v) = 1$ for all $j \in \{m, \ldots, m'\}$. Hence the invocation of ReadMem in Line (6) returns a value other than $K$ and $P_i$ executes Line (8). It remains to show Jump returns $r_{\max}$ for the round number.

Note that Jump returns $r$ if $r \geq R$, in which case $r = R = r_{\max}$. Otherwise, let $c$ denote $[\log(R - r)]$. The while-loop of Jump calculates the following sequence $\{r'_0, \ldots, r'_c\}$ of natural numbers: $r'_0$ is $r$ and $r'_{i+1}$ is

- $r'_i$, if $r'_i + 2^{c-i} > R$ or ReadMem($r'_i + 2^{c-i}, K$) returns $K$;
- $r'_{i+1} := r'_i + 2^{c-i}$, otherwise.
From this we obtain a nested sequence of intervals:

\[
[r'_0, r'_0 + 2^c), \ldots, [r'_i, r'_i + 2^{c-i}), \ldots, [r'_c, r'_c + 2^0).
\]

It is easy to see that \( r_{\max} \) belongs to every one of these intervals and, since the last is a singleton, we know \( r'_c = r_{\max} \). This is precisely the round number returned by Jump.

**Lemma 11.** Suppose ModifiedCIL starts from a reachable state \( s \) and suppose \( r \leq R - 2 \). Then, with probability greater than 0.511, at least one process terminates successfully in a round no higher than \( r + 2 \). Moreover, at most \( 15N \) complete phases are executed between \( s \) and the successful termination.

**Proof.** By assumption, at least one process survives throughout the execution of the algorithm. Therefore, if no successful termination ever takes place, the algorithm stops only if all surviving processes reach round \( R \) and switch to the exit algorithm. Without loss of generality, we assume that no successful termination occurs in round \( r + 1 \) or lower.

Consider the first complete phase following \( s \). There are two cases.

- It is executed by a process \( P_1 \) in round \( < r \). By Lemma 10, \( P_1 \) reaches round \( r \) by the end of this phase.
- It is executed by a process \( P_1 \) in round \( r \). Then \( P_1 \) reaches Line (7) in this phase and, with probability \( \frac{1}{2N} \), \( P_1 \) advances to round \( r + 1 \).

Suppose that either the first case applies, or the second case applies but \( P_1 \) fails to advance to round \( r + 1 \). Then the same case distinction can be made on the next complete phase. This repeats until all surviving processes have reached round \( r \) and, after that point, every complete phase involves a coin toss to advance to round \( r + 1 \) until a success occurs. Moreover, since a lagging process catches up to round \( r \) in one complete phase, at most \( N \) complete phases following \( s \) are executed by processes in round strictly lower than \( r \).

Consider the event \( E_1 \), in which “a success occurs before \( 5N \) attempts to move from \( r \) to \( r + 1 \) are made” and “all subsequent attempts to move from \( r \) to \( r + 1 \) fail.” Notice the first condition is equivalent to “it is not the case that all of the first \( 5N \) attempts to move from \( r \) to \( r + 1 \) fail,” which occurs with probability \( 1 - (1 - \frac{1}{2N})^{5N} \). By the reasoning above, this first success occurs before \( 6N \) complete phases are executed following \( s \).

Let \( P_1 \) be the successful process, thus the first to reach round \( r + 1 \). By our atomicity assumption, \( \text{mem}(r + 1, s, p_i) \) is set to 1 as soon as \( P_1 \) reaches round \( r + 1 \). After that point, every other \( P_v \) tosses a coin at most once to advance from \( r \) to \( r + 1 \). This is because in the subsequence phase \( P_v \) sees it’s no longer leading and therefore does not execute Line (7). As a result, the probability of “all subsequent attempts to move from \( r \) to \( r + 1 \) fail" is at least \( (1 - \frac{1}{2N})^{5N-1} \) and hence the probability of \( E_1 \) is at least \( 1 - (1 - \frac{1}{2N})^{5N} ) (1 - \frac{1}{2N})^{N-1} \). Moreover, after \( P_1 \) reaches round \( r + 1 \), at most \( 2N - 2 \) complete phases are executed by processes in round \( r \) or lower; at most \( N - 1 \) failed coin tosses to move from \( r \) to \( r + 1 \) and at most \( N - 1 \) phases to catch up to \( r + 1 \).
By assumption, no successful termination takes place until a process has reached round \( r + 2 \). Thus, every complete phase executed by a process in round \( r + 1 \) is a coin toss to move to round \( r + 2 \), until a success occurs. Let \( E_2 \) denote the event that “a success occurs before 5N attempts to move from \( r + 1 \) to \( r + 2 \) are made.” The probability of \( E_2 \) given \( E_1 \) is then \( 1 - (1 - \frac{1}{2N})^{5N} \). Similarly, this success occurs before \( 6N + (2N - 2) + 5N = 13N - 2 \) complete phases are executed following \( s \) and, after this success, at most 2N - 2 complete phases are executed by processes in round \( r + 1 \) or lower.

Therefore, given \( E_1 \) and \( E_2 \), at least one process executes a complete phase in round \( r + 2 \) before 15N complete phases are executed following \( s \). Due to \( E_1 \), no process reaches round \( r + 1 \) with preference value other than \( s, p_i \). Therefore the first process to complete a phase in round \( r + 2 \) sees no disagreement in round \( r + 1 \) or higher. It then terminates successfully by Line (4). It remains to consider the probability of both \( E_1 \) and \( E_2 \) occurring. Recall that \( \{ (1 - \frac{1}{n})^n \}_{n=2}^\infty \) is an increasing sequence with limit \( \frac{1}{e} \). Therefore \( \{ (1 - \frac{1}{2n})^{2n} \}_{n=2}^\infty \) is a decreasing sequence with limit \( \frac{1}{e^2} \) and \( \{ (1 - \frac{1}{2n})^{2n} \}_{n=2}^\infty \) is a decreasing sequence with limit \( 1 - \frac{1}{e^2} \). Therefore, the probability of both \( E_1 \) and \( E_2 \) occurring is at least

\[
(1 - (1 - \frac{1}{2N})^{5N})^2 \cdot (1 - \frac{1}{2N})^{N-1} \geq (1 - \frac{1}{e^{2.5}})^2 \cdot \frac{1}{\sqrt{e}} > 0.511.
\]

Notice Lemma 11 applies only to executions starting in round \( R - 2 \) or lower. The next lemma covers rounds \( R - 1 \) and \( R \), assuming a decision is reached before switching to the exit algorithm.

**Lemma 12.** Suppose ModifiedCIL starts from a reachable state \( s \). Let \( r \) denote \( s, r_{\text{max}} \) and suppose \( R - 2 < r \leq R \). Assuming the exit algorithm is not invoked, the (conditional) probability that at least one process terminates successfully before 15N complete phases are executed after \( s \) is greater than 0.511.

**Proof.** We use arguments similar to those in the proof of Lemma 11. First suppose \( r = R \). Then at most \( N - 1 \) complete phases are executed before a process completes a phase in round \( R \). Suppose \( P_i \) is the first to do so. If \( P_i \) does not terminate by Line (4) in that phase, it must be the case that \( \text{mem}(R - 1, v_1) = \text{mem}(R - 1, v_2) = 1 \) for some \( v_1 \neq v_2 \). Then \( P_i \), as well as every other process that reaches round \( R \), invokes the exit algorithm. By assumption, this is not the case and hence \( P_i \) terminates by Line (4) in that phase. Therefore, with probability 1, at least one process terminates before \( N \) complete phases are executed.

If \( r = R - 1 \), then at most \( N - 1 \) complete phases are executed before a process completes a phase in round \( R - 1 \). Similar to the previous case (\( r = R \)), if the first process completing a phase in round \( R - 1 \) does not terminate by Line (4) in that phase, every process reaching round \( R - 1 \) will try to advance to round \( R \) by Line (7), until one of them succeeds. The probability of at least one success before \( 4N \) attempts are made is \( 1 - (1 - \frac{1}{2N})^{4N} \), which is bounded
below by \((1 - \frac{1}{r}) > 0.864\). After that success, the problem reduces to the case where \(r = R\) and successful termination is guaranteed before \(N\) complete phases. Therefore, with probability at least 0.864, at least one process terminates before 6\(N\) complete phases are executed.

**Theorem 3.** If the exit algorithm is wait-free and satisfies probabilistic termination, the same holds for ModifiedCIL.

**Proof.** By correctness of the exit algorithm, we may focus on the case in which the exit algorithm is not invoked. Consider execution blocks of 15\(N\) complete phases each. By Lemma 11 and Lemma 12, the probability of successful termination within each block is at least 0.511. Thus, with probability 1, the algorithm terminates successfully after a finite number of blocks. Since we have made no assumption on the number of surviving processes, the algorithm is wait-free.

We now turn to complexity considerations. Again, we make a case distinction based on whether the exit algorithm is invoked.

**Lemma 13.** Assume that the exit algorithm is not invoked. The expected number of elementary read/write operations until at least one process terminates successfully is \(O(N \log(\log N))\).

**Proof.** Again we consider execution blocks of 15\(N\) complete phases each. The expected number of blocks is:

\[
\sum_{n=1}^{\infty} (n \cdot 0.511 \cdot (1 - 0.511)^{n-1})
\]

\[
= \sum_{n=0}^{\infty} ((n + 1) \cdot 0.511 \cdot 0.489^n)
\]

\[
= 0.511 \cdot \left( \sum_{n=0}^{\infty} 0.489^n + \sum_{n=1}^{\infty} 0.489^n + \sum_{n=2}^{\infty} 0.489^n + \ldots \right)
\]

\[
= 0.511 \cdot \left( \sum_{n=0}^{\infty} 0.489^n + 0.489 \sum_{n=1}^{\infty} 0.489^n + 0.489^2 \sum_{n=0}^{\infty} 0.489^n + \ldots \right)
\]

\[
= 0.511 \cdot \left( \sum_{n=0}^{\infty} 0.489^n \right)^2
\]

\[
= 0.511 \cdot \left( \frac{1}{0.511} \right)^2 = \frac{1}{0.511} < 2.
\]

Thus the expected number of complete phases is at most 30\(N\). Moreover, there are at most \(N - 1\) incomplete phases. Since each phase involves \(O(\log(\log N))\) elementary operations, the expected number of elementary operations is at most \(O(N \log(\log N))\).

**Lemma 14.** Suppose the exit algorithm is the original CIL algorithm and is invoked. The expected number of elementary read/write operations until at least one process terminates successfully is \(O(N^2 \log(\log N))\).
Proof. In this case the algorithm steps through all \( R \) rounds without successful termination. Using a similar calculation as in the proof of Lemma 13, the expected number of coin tosses to move from \( r \) to \( r + 1 \) is
\[
\sum_{n=1}^{\infty} n \left( \frac{1}{2N} \right) (1 - \frac{1}{2N})^{n-1} = 2N.
\]
Following each success, at most \( N - 1 \) phases are executed by processes lagging behind. Therefore, the expected number of complete phases before switching to original CIL is at most \( 3NR \leq 6N(\log N + 1) \). The expected number of elementary operations before switching is then \( O(N(\log N)(\log(\log N))) \).

In [CIL94], it is shown that the expected number of elementary operations for the original CIL algorithm is \( O(N^2) \). Therefore, the overall expected number of elementary operations is \( O(N^2 \log(\log N)) \). \( \square \)

**Lemma 15.** Suppose the ModifiedCIL starts from the initial state \( s_0 \). The probability of failing to reach a decision in or before round \( R \) is at most \( 1/N \).

Proof. By Lemma 11, this probability is at most \( (1 - 0.511)^{\frac{R}{2}} \). Since \( R = 2[\log N] \), we have
\[
(1 - 0.511)^{\frac{R}{2}} \leq (1 - 0.511)^{\log N} < (0.5)^{\log N} = \frac{1}{N}.
\]
\( \square \)

Putting together Lemmas 13, 14, and 15, we conclude that the expected complexity of ModifiedCIL is \( O(N \log(\log N)) \).

**Theorem 4.** Suppose ModifiedCIL starts from the initial state \( s_0 \) and the exit algorithm is original CIL. The expected number of elementary read/write operations until at least one process terminates successfully is \( O(N \log(\log N)) \).

### 6 Model Checking

It turns out to be quite straightforward to specify our algorithm in PRISM’s state-based input language. Each process is modeled as a *module* and the shared memory is modeled using global variables. Two more global variables are used to keep track of process failures and the number of completed phases.

We consider binary consensus (i.e., \( K = 2 \)) with \( N = 2, 3, 4 \) processes. Processes are assumed to disagree initially, therefore validity is trivial. Agreement is satisfied in all models constructed. For probabilistic termination, we ask PRISM to compute the (exact) minimum probability of at least one process terminating successfully, given an allowance of \( R = 2[\log N] \) rounds and \( 15N \cdot \frac{R}{2} = 15N[\log N] \) complete phases. This result is compared against our analytic lower bound of \( 1 - \frac{1}{N} \).

In the case of \( N = 4 \), the model becomes too complex (with \( 2[\log N] = 4 \) rounds and \( 15N[\log N] = 120 \) complete phases). However, we discover that the
analytic bound of $1 - \frac{1}{N} = 0.750$ is already met when we restrict to 40 complete phases. This suggests that we have made some overly conservative estimates while deriving the analytic bound.

The table below summarizes our results. We use PRISM version 2.1, running on a 1.4 GHz Pentium M machine with 500 Mb memory under Linux 2.6. The MTBDD engine is used with a CUDD memory limit of 400 Mb. Other parameters remain at default settings. All relevant files, including model checking logs, can be found in [Che05].

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7 Conclusions

We have given a relatively simple algorithm that solves asynchronous wait-free consensus in expected $O(N \log(\log N))$ total work. To our best knowledge, this is the most efficient algorithm (in terms of expected total work) for dynamic adversaries. Moreover, our atomicity assumption is more benign than common assumptions such as value-oblivious and write-oblivious, making our algorithm more widely applicable.

We make use of MWMR memory in order to reduce global and local data. This strategy, also adopted in [Cha96, Aum97], leads to more efficient consensus algorithms. Interestingly, it also makes model checking significantly more feasible, for it helps to avoid the typical state explosion problem. MWMR memory is often regarded as a stronger primitive than SWMR memory. Indeed, there are optimal implementations of MWMR from SWMR physical registers using linear time and logarithmic space [IS92]. However, dissenting opinions appear in [BPSV00], where the authors argue that SWMR memory requires the hidden assumption of naming: existence of distinct identifiers known to all. In that sense, MWMR is weaker than SWMR.

In theory, we can implement our MWMR algorithm using physical SWMR registers via an $O(N)$ emulation of MWMR from SWMR. Then the total number of physical memory operations is raised by a factor of $N$ to $O(N^2 \log(\log N))$, but the total number of logical memory operations remain at $O(N \log(\log N))$. This property can be quite useful in reducing network communication costs, in case remote processes participate in the consensus protocol. In other words, one can aim at an implementation in which the traffic to and from a remote participant corresponds to the number of logical operations performed by that participant, as opposed to the much higher number of physical operations.

For future work, we want to improve the per process work bound of our algorithm. In [AW96], a similar improvement is achieved by allowing fast processes to cast votes of increasing weights. However, their proofs rely on properties of
Martingale processes, which are not directly applicable to our algorithm. It is also possible that in our setting per process work is inherently high, e.g. \( \Omega\left(\frac{N}{\log N}\right) \), where \( f \) is a polylogarithmic function.

Another possibility for future work is to consider contention cost, which measures the amount of conflict in memory access [AB04]. The contention cost for ModifiedCIL is high because, in a roughly synchronous execution, all \( N \) processes try to access a constant number of registers at the same time. It would be interesting to modify the algorithm further to reduce contention.

Finally, we comment on model checking using PRISM. Although the current limit seems to be 4 processes, we conjecture a vast improvement using a symmetry reduction option, which is under development by the PRISM team. Before symmetry reduction is available, manual abstraction can be used to increase feasibility. That is, we manually construct an abstraction that captures core ideas of an algorithm, while significantly decreasing the model size. We experimented with such an abstraction of original CIL, by focusing on the shared memory and filtering out local states of processes. Having done so, we were in fact able to handle up to 10 processes. However, it is non-trivial to prove soundness of the abstraction. Standard techniques such as probabilistic simulation are available for this purpose, but substantial investment of time is required.

Overall, PRISM allows us to conduct experiments during the development stage of an algorithm, with minimal learning effort. Although in most cases it still cannot handle large instances of a full algorithm, it is perfectly feasible to model check a subroutine or an abstract version. This already provides valuable information, especially to those who simply wish to gain more insight into an algorithm.

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References


