Revisiting the Complexity of Hardware Cache Coherence and Some Implications

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Cache coherence is an integral part of shared-memory systems but is also widely considered to be one of the most complex parts of such systems. Much prior work has addressed this complexity and the verification techniques to prove the correctness of hardware coherence. Given the new multicore era with increasing number of cores, there is a renewed debate about whether the complexity of hardware coherence has been tamed or whether it should be abandoned in favor of software coherence. This article revisits the complexity of hardware cache coherence by verifying a publicly available, state-of-the-art implementation of the widely used MESI protocol, using the Mur ϕ model checking tool. To our surprise, we found six bugs in this protocol, most of which were hard to analyze and took several days to fix. To compare the complexity, we also verified the recently proposed DeNovo protocol, which exploits disciplined software programming models. We found three relatively easy to fix bugs in this less mature protocol. After fixing these bugs, our verification experiments showed that, compared to DeNovo, MESI had 15X more reachable states leading to a 20X increase in verification (model checking) time. Although we were eventually successful in verifying the protocols, the tool required making several simplifying assumptions (e.g., two cores, one address). Our results have several implications: (1) they indicate that hardware coherence protocols remain complex; (2) they reinforce the need for protocol designers to embrace formal verification tools to demonstrate correctness of new protocols and extensions; (3) they reinforce the need for formal verification tools that are both scalable and usable by non-expert; and (4) they show that a system based on hardware-software co-design can offer a simpler approach for cache coherence, thus reducing the overall verification effort and allowing verification of more detailed models and protocol extensions that are otherwise limited by computing resources.

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1. INTRODUCTION

Shared memory continues to be one of the most widely used parallel programming models due to the advantages offered by a global address space. Today's multicores directly support shared memory and form the heart of a variety of computing systems, ranging from smartphones to data centers to the largest supercomputers. One of the main challenges of shared memory hardware is to keep caches coherent, and this is usually addressed by hardware cache coherence protocols. Traditionally, a major problem with such protocols has been that they have been extremely complex, and implementing bug-free protocols has been a challenge [Abts et al. 2003]. Increasing on-chip core counts in manycores exacerbates this problem because it demands moving to variants of directory-based protocols. Methods like trace-driven or other simulation-based testing are insufficient for ensuring the correctness of such complex protocols since they often do not hit the difficult corner cases. Formal verification techniques have therefore been proposed and deployed to prove the correctness of these protocols [Abts et al. 2000; Burckhardt et al. 2005; Gjessing et al. 1989; McMillan and J. 1991; Nanda and Bhuyan 1992; Park and Dill 1995; Pong et al. 1998].

Over the past several decades, there has been much experience accumulated in both designing cache coherence protocols and building formal methods based tools to verify them. Nevertheless, as the number of cores on a chip grows, the power, performance, and area inefficiencies of hardware cache coherence (e.g., directory storage, indirection through the directory, invalidation and acknowledgment messages, false sharing, cache line based granularities, and so on) are potential obstacles to scalability. To reduce these inefficiencies, researchers continue to propose enhancements to basic coherence protocols—these enhancements typically come at the cost of even further complexity [Zebchuk et al. 2009; Zhao et al. 2013].

There is a debate in the community about whether we have tamed the complexity of hardware cache coherence enough to continue building increasingly larger hardware cache coherent systems or whether we must abandon hardware coherence in favor of other alternatives. Thus, on the one hand, companies like Tilera have built hardware cache coherent systems with as many as 64 cores, making the resulting familiar programming model a key differentiator for their product line. A recent position paper makes the case that hardware coherence is here to stay [Martin et al. 2012]. Other researchers continue to extend current protocols in new ways that make them even more complex [Aisopos and Peh 2011]. On the other hand, Intel has recently built a 48-core Single-Chip Cloud Computer (SCC) [IntelSCC 2009] that abandons hardware cache coherence, replacing it with software managed caches and their attendant programming complexity. Academic and industry researchers have also proposed systems that eliminate dependence on hardware coherence in various ways [Borkar 2011; Choi et al. 2011; Kelm et al. 2009; Lin et al. 2012].

In this article, we revisit the complexity of hardware cache coherence by asking the following questions. Has our experience with hardware cache coherence in the past few decades enabled us to conquer its complexity? If not, what are the implications for future system design?

To answer these questions, ideally, we would use an industry-strength cache coherence protocol specification and quantitative metrics for its complexity. Given the proprietary nature of industry protocols, we instead use a state-of-the-art, mature, publicly available protocol, the MESI protocol implemented in the Wisconsin GEMS simulation

suite (version 2.1.1) [Martin et al. 2005].¹ GEMS was released in 2005 and remains a widely used simulation package in the computer architecture community. Unfortunately, however, there are no currently accepted quantitative metrics to measure complexity. Instead, we report on our experiences verifying the GEMS MESI protocol using the Mur ϕ model checking tool (version 3.1, slightly modified to exploit 64-bit machines) [Dill et al. 1992; Ip and Dill 1993, 1996]. Although more advanced verification techniques exist, we chose Mur ϕ for its easy-to-use interface and robustness. Mur ϕ has also been the tool of choice for many hardware cache-related studies [Abts et al. 2000; Burckhardt et al. 2005; Park and Dill 1995; Pong et al. 1998; Zhang et al. 2010].

Our results from applying Mur ϕ to the GEMS MESI protocol were surprising. We found six bugs in the protocol, even though it is a mature protocol used by a large number of architecture researchers. More significantly, some of these bugs involved subtle races and took several days to debug and fix. We contacted the developers of the GEMS simulation team with our bug findings in 2011. They had seen one of the six bugs, bug #4 in Section 4.1, but were surprised by the other bugs. Some of these bugs were also present in the GEM5 simulator [Binkert et al. 2011], an extension to the GEMS simulator that incorporates the M5 CPU core simulator, at that time. After we showed our fixes, the GEMS group fixed the bugs and released new patches. These fixes needed the addition of multiple new state transitions and extra buffer space for stalling requests in the protocol. After fixing the protocol, we were able to successfully run Mur ϕ , but the tool required making several simplifying assumptions typical of such tools today (e.g., two cores, single cache line per core, and one address). With these assumptions, the verification process explored 1,257,500 reachable states and took 173 seconds using the default breadth-first search option of Mur ϕ . We ran out of memory on our machine (with 32GB of RAM) when we tried to model a system containing two memory addresses.

Our conclusion from this experience is that hardware cache coherence protocols still remain difficult to design, understand, and prove correct. Although formal verification techniques have been embraced by industry, their use is not yet common practice for architecture researchers. Our results reinforce the need to use formal verification tools to assess the correctness and complexity of coherence protocols. Conversely, our work also motivates the need for verification tools that are scalable to larger systems and usable by non-expert.

An extreme solution to the hardware complexity problem is represented by the Intel SCC [IntelSCC 2009]. The SCC eliminates the hardware complexity of cache coherence but imposes significant programming complexity due to the requirement of managing the caches in software. Recently, we proposed an alternate method, called DeNovo, that is based on a tighter hardware/software co-design approach [Choi et al. 2011]. For easier programmability, this approach espouses disciplined shared-memory programming languages that eliminate data races by design, provide determinism-by-default, and provide safe nondeterminism only when explicitly requested. With a guarantee of data-race-freedom from software, DeNovo is able to avoid races in hardware and the resulting transient states and their complexity. The simplicity of the protocol, coupled with the disciplined software, provides performance and power advantages to DeNovo as well; furthermore, various enhancements (optimizations) can be incorporated without adding new states or transitions [Choi et al. 2011]. We report our experiences with using Mur ϕ to verify the DeNovo protocol and compare them with the MESI experiences.

To verify DeNovo, we first extended the canonical coherence protocol modeling technique in Mur ϕ to model the software constraints that DeNovo hardware expects. In

¹We chose MESI over a more complicated MOESI from GEMS because the GEMS web site says that the CMP version of the latter has bugs. Although we specifically study the MESI protocol in this article, we believe our insights apply to other protocols, such as MOESI, as well.

particular, this extension is able to guarantee the data-race-free property and a disciplined parallel phase behavior. Using this extension, we verified the base DeNovo protocol. Despite DeNovo's immaturity, we found only three bugs in the implementation. Furthermore, these bugs were simple to fix and turned out to be mistakes in translating the high level description of the protocol into the implementation (i.e., their solutions were already present in the internal high level description of the protocol). After fixing the bugs, we found that, compared to MESI, Mur φ had to explore 15X fewer reachable states for DeNovo, taking 20X less time to fully model check DeNovo. It is difficult to define a metric to quantify the relative verification complexity of coherence protocols; nevertheless, our results demonstrate that hardware-software co-designed approaches like DeNovo can lead to much simpler protocols than conventional hardware cache coherence (while providing an easy programming model, extensibility, and competitive or better performance [Choi et al. 2011]).²

Overall, this article shows that despite the accumulated experience in designing hardware cache coherence, these protocols remain complex. It reinforces the need to use formal verification by protocol designers, motivating scalable and easy-to-use tools for non-experts. Finally, the article motivates hardware-software co-designed protocols that are much simpler to design and verify than pure hardware coherence and that exploit simpler programming models.

2. COMPLEXITY OF CACHE COHERENCE

Cache coherence is usually defined as a mechanism that ensures writes to the same memory location appear in the same order to all cores. In this section, we discuss two directory based protocols, the MESI protocol, which is based on a state-of-the-art implementation obtained from the publicly available GEMS simulator [Martin et al. 2005], and the recent DeNovo protocol [Choi et al. 2011], which is based on our own implementation. Without loss of generality, we assume a multicore system with n cores, private $L1$ caches, a shared $L2$ cache, and a general (nonbus, unordered) interconnect on chip.

2.1. MESI

MESI, also known as the Illinois protocol [Papamarcos and Patel 1984], stands for *Modified* (locally modified and no other cache has a copy), *Exclusive* (unmodified and no other cache has a copy), *Shared* (unmodified and some other caches may have a copy), and *Invalid*. Over the MSI protocol, the *Exclusive* state has the added advantage of avoiding invalidation traffic on write hits. For scalability, we assume a directory protocol [Lenoski et al. 1992]. Given our shared (inclusive) $L2$ cache-based multicore, we assume a directory entry per $L2$ cache line, referred to as an in-cache directory [Censier and Feautrier 1978]. We use $L2$ and *directory* interchangeably.

Figure 1 shows the simple textbook state transition diagram for an $L1$ cache with the MESI protocol. The $L2$ cache also has four (textbook) states: $L1_Modified$ (modified in a local $L1$), $L2_Modified$ (modified at $L2$ and not present in any $L1$), $Shared$ (valid data at $L2$ and present in one or more $L1$ s), and $Invalid$. It also has a *dirty* bit (set on receiving a writeback from $L1$) that indicates whether data is dirty or not. When in $Shared$ state, the $L2$ /directory contains the up-to-date copy of the cache block data along with a list of sharers. On a read miss request, the *directory* services the request

²The paper proposing DeNovo [Choi et al. 2011] showed that the basic DeNovo protocol is performance-competitive with MESI, and its simplicity enables optimizations that significantly improved performance and network traffic (and thereby energy) [Choi et al. 2011]. That paper focused on performance, but as evidence of DeNovo's simplicity, it also briefly summarized some of the key results of this article. That paper only gave the number of bugs found in each protocol, the number of states explored, and the verification times. It did not describe the nature of the bugs in MESI or DeNovo, the fixes involved, or the modification to standard Mur φ modeling required to verify a hardware/software co-designed protocol like DeNovo.

Fig. 2. Example state transitions for MESI.

transition tables shown in Figure 2(c) and (d). *Ack_all* is triggered for the last incoming *Ack* and *Data_Ack_all* is triggered if L2's data response (which includes the *Ack* count) is the last message to be received.

This example illustrates the need for transient states and the additional complexities introduced by them in the MESI protocol. The larger the number of transient states, the more complex the protocol becomes. In the GEMS implementation of the MESI protocol, there are 7 transient states in L1 and 14 transient states in L2. Optimizations to the described baseline protocol usually incur additional transient states.

2.2. DeNovo

One of the main difficulties in writing and debugging parallel programs is data races. These software data races are also responsible for races in hardware with conventional hardware cache coherence protocols. These hardware races are directly responsible for most transient states in hardware protocols, and they form the root cause of much of the complexity in such protocols.

There has been a lot of software work on disciplined shared-memory programming models [Anderson et al. 2008; Gustafsson 2009; Berger et al. 2009; Blumofe et al. 1995; Budimlic et al. 2009; Ghuloum et al. 2007; Bocchino et al. 2011, 2009; Olszewski et al. 2009; Kulkarni et al. 2007; Allen et al. 2009] to eliminate or alleviate the impact of data races, with the goal of providing better programmability, debuggability, and concurrency safety of parallel programs. Along with addressing the data race problem, much of this work also provides strong safety properties, such as determinism and various disciplined forms of nondeterminism, to *simplify parallel programming*. DeNovo is a hardware architecture that exploits the properties of such disciplined software

programming models to additionally *simplify hardware* while also providing performance, power, and scalability benefits [Choi et al. 2011].

2.2.1. Software Model. DeNovo currently uses Deterministic Parallel Java (DPJ) [Bocchino et al. 2011, 2009] as an example of a disciplined programming language to drive its design. DPJ is an extension to Java that provides properties like structured parallel control, data-race-freedom, and explicit specification of the effects of shared memory accesses. In a DPJ program, the programmer partitions the heap into named “regions” and annotates methods with summaries of read and write operations on these regions (“effects”). The DPJ compiler uses this region and effect information to guarantee determinism.

This article does not quantify software complexity; however, our software philosophy and DPJ are motivated entirely by the goal of *reducing software complexity*. Even today, the C++ [Boehm and Adve 2008] and Java [Manson et al. 2005] memory models do not provide any reasonable semantics for data races; therefore, a data race in these programs is a bug and imposes significant verification complexity. In contrast, DPJ provides strong safety guarantees of data-race-freedom and determinism-by-default. Programmers can reason about deterministic programs as if they were sequential. There is certainly an additional up-front burden of writing region and effect annotations in DPJ; however, arguably, this burden is mitigated by the lower debugging and testing time afforded by deterministic-by-default semantics. There is also ongoing work on partly automating the insertion of these annotations [Vakilian et al. 2009]. Overall, there is a growing part of the community that is working toward similar goals of more disciplined programming practices [WoDET 2013].

In this article, we focus on the baseline DeNovo protocol [Choi et al. 2011] targeted to deterministic codes based on the original DPJ language proposed in Bocchino et al. [2009]. DPJ and DeNovo are ongoing projects that continue to incorporate support for increasingly wider classes of programs [Bocchino et al. 2011; Bocchino and Adve 2011; Heumann et al. 2013; Sung et al. 2013, 2014]. We discuss this ongoing work and implications for general codes in Section 5.

2.2.2. DeNovo Hardware. DeNovo exploits disciplined language features for a more efficient, scalable, and simpler hardware. Specifically, structured parallel control and explicit read and write effect information enable caches to self-invalidate stale data (communicated through a compiler-inserted self-invalidate instruction for each region with write effects) and thus remove the need to maintain sharer lists and to send invalidation and acknowledgment messages. In addition, the data-race-freedom guarantee makes it possible to completely eliminate transient states from the cache coherence protocol and thus greatly reduce its complexity. As shown in [Choi et al. 2011], these benefits do not compromise the performance of the DeNovo protocol; instead, they provide opportunities for improved performance and energy without increasing complexity.

The DeNovo protocol has three states at *L1*, *Registered* (locally modified), *Valid*, and *Invalid*. *L2* has the same states, too, with obvious meaning. The handling of read requests is straightforward. On a write, if the line is not registered, *L1* issues a registration request to *L2* and transitions to *Registered* state immediately because there are no races in the program. Similar to MESI, DeNovo also has a *dirty* bit at *L2*, which is set to indicate dirty data. The complete state transition table and a detailed description of the DeNovo protocol can be found in Choi et al. [2011].

Let us revisit the code segment from Figure 2. Figure 3(a) shows the changes to the code required to prove data-race-freedom. Specifically, the shared variable *A* is placed in a region *R_A*, both the parallel phases are annotated with read and write effect summaries, and, finally, a self-invalidation instruction is inserted at the end of the second phase.

Fig. 3. Example state transitions for DeNovo.

Figure 3(b) shows the timeline of the state transitions for the DeNovo protocol and the state transition tables for the states encountered in this example are shown in Figure 3(c) and (d). Focusing again on the write instruction in the second phase, $L1_{P1}$ transitions directly to the *Registered* state without transitioning to any transient state and sends a registration request to $L2$. $L2$, on receiving the registration request, transitions to the *Registered* state. We do not show the registration response message from $L2$ here because it is not in the critical path and is handled by the request buffer at $L1$. As explained in Choi et al. [2011], the functionality of the request buffer is similar to that of a uniprocessor system and is invisible to the outside world. At the end of the phase, each core executes a self-invalidate instruction on region R_A in the $L1$ cache of its core (except for data in *Registered* state since this data is known to be up-to-date).

As an optimization, the $L1$ maintains an additional *touched* bit for each word. This bit is set on a read, indicating that this data will stay up-to-date until the end of the phase. This is guaranteed by data-race-freedom because there cannot be concurrent writers to the same location. The self-invalidation operation now skips invalidating the data that is both *Valid* and *touched*. These *touched* bits are reset at the end of the parallel phase.

This example illustrates how the absence of transient states makes the DeNovo protocol simpler than MESI.

3. MODELING FOR PROTOCOL VERIFICATION

We use the Mur ϕ model checking tool [Dill et al. 1992; Ip and Dill 1993, 1996] to verify the simple word-based protocols (equal address, communication and coherence granularity as explained in Choi et al. [2011]) of DeNovo and MESI. We derived the MESI model from the GEMS implementation [Martin et al. 2005]. We derived the DeNovo model from our implementation. To keep the number of states explored (by Mur ϕ) tractable, as is common practice, we used a single address, single region (only for DeNovo), two data values, and two cores. We modeled private $L1$ caches, a unified

Fig. 4. State transitions for *AccessStatus* data structure in a given phase. An access for which there is no transition cannot occur and is dropped.

L2, an in-cache directory (for MESI), and an unordered full network with separate request and reply links. Both models allow only one request per *L1* in the rest of the memory hierarchy. Because we modeled only one address, we modeled replacements as unconditional events that can be triggered at any time. To enable interactions across multiple parallel phases (cross-phase) in both the models, we introduced the notion of a phase boundary by modeling it as a sense-reversing barrier. Finally, we modeled the data-race-free guarantee for DeNovo by limiting conflicting accesses. We explain each of these attributes in detail in the following sections.

3.1. Abstract Model

To reduce the amount of time and memory used in verification, we modeled the processors, addresses, data values, and regions as *scalarsets* [Ip and Dill 1996], a datatype in *Murφ* that takes advantage of the symmetry in these entities while exploring the reachable states. A processor is modeled as an array of cache entries consisting of *L1* state information along with protocol-specific fields like the region field and the *touched* bit for DeNovo. *L1* state is one of three possible states for DeNovo or one of 11 possible states for MESI. Similarly, *L2* is also modeled as an array of cache entries, each with *L2* state information, a *dirty* bit, and other protocol-specific details like sharer lists for MESI. *L2* state is one of three possible states for DeNovo or one of 18 possible states for MESI. Memory is modeled as an array of addresses storing data values.

3.1.1. Data-Race-Free Guarantee for DeNovo. To model the data-race-free guarantee from software for DeNovo, we used an additional data structure called *AccessStatus*. As shown in Figure 4, this data structure maintains the current status (*read*, *readshared*, or *written*) and the core id of the last requestor for every address in the model. The current status and the last requestor determine the reads and writes that cannot occur in a data-race-free program and are thus disallowed in the model.

On any read, if it is the first access to this address in this phase, then *status* is set to *read*. If *status* is already set to *read* and the requesting core is not the same as the last requestor, then *status* is set to *readshared*. If *status* is *readshared*, then it stays the same on the read. If *status* is *written* and the requesting core is the same as the last requestor, it stays as *written*. On the other hand, if the requesting core is not the same

as the last requestor, then this access is not generated in the model since it violates the data-race-freedom guarantee.

Similarly, on any write, if it is the first access to this address or if the requesting core is the same as the last requestor, then *status* is set to *write*. If *status* is either *readshared* or the requesting core is not the same as the last requestor, then this access is not generated to adhere to the data-race-free guarantee.

The *AccessStatus* data structure is reset for all the addresses at the end of a phase.

3.1.2. Cross-phase Interactions. We modeled the end of a parallel phase (and the start of the next phase) using a sense-reversing barrier implementation [Mellor-Crummey and Scott 1991]. This event (end-of-phase) can be triggered at any time; that is, with no condition. The occurrence of end-of-phase is captured by a flag, *releaseflag*. This event occurs per core and stalls the core from issuing any more memory requests until (1) all the pending requests of this core are completed (i.e., the *L1* request buffer is empty), and (2) all other cores reach the barrier. The completion of end-of-phase is indicated by resetting the *releaseflag* flag. Figure 5 shows the Mur ϕ code for end-of-phase implementation for the DeNovo protocol. The *spinwaiting* flag indicates that the current core is waiting for other cores to reach the barrier. When a core enters the barrier for the first time, the local sense of the barrier (*localsense*) is reversed, thus indicating entering a new barrier; barrier count (*barcount*) is updated; and the *spinwaiting* flag is set. If it is the last one to enter the barrier, the core also notifies all other cores about the end of barrier by assigning *barrier* its *localsense*. It also resets the *barcount* and *releaseflag*. Once a core reaches the barrier, we also modeled self-invalidations and unsetting *touched* bits for DeNovo. The code for MESI is similar except for DeNovo-specific operations like self-invalidation and unsetting *touched* bits.

3.2. Invariants

This section discusses the invariants we checked to verify the MESI and DeNovo protocols. The MESI invariants are based on prior work in verification of cache coherence protocols [Dill et al. 1992; McMillan and J. 1991]. The DeNovo invariants are analogous as further described later. (Adding more invariants does not affect the verification time appreciably because the number of system states explored is still the same.)

3.2.1. MESI Invariants. We used five invariants to verify the MESI protocol [Dill et al. 1992; McMillan and Schwalbe 1991].

- Empty sharer list in Invalid state.* This invariant asserts that the sharer list is empty when *L2* transitions to *Invalid* state, and it ensures that there are no *L1*s sharing the line after *L2* replaces the line.
- Empty sharer list in Modified state.* This invariant asserts that the sharer list is empty when *L2* transitions to *Modified* state.
- Only one modifiable or exclusive cache copy.* This invariant checks that there is only one cache copy in either *Modified* or *Exclusive* state. It is also a violation for a cache line to be in both these states at the same time.
- Data value consistency at L1.* When *L1* is in *Shared* state and *L2* is also in *Shared* state, the data values should be the same at both *L1* and *L2*. Indirectly, this invariant also makes sure that all the *L1*s have the same data value when in *Shared* state.
- Data value consistency at L2.* This invariant checks that when *L2* is in *Shared* state and *dirty* bit is not set, *L2*'s data value should be the same as at memory.

3.2.2. DeNovo Invariants. We modeled six invariants for the DeNovo protocol. Because there is no sharer list maintained in the DeNovo protocol, we do not check for the first two invariants of the MESI protocol. The first three invariants of the DeNovo protocol

```

/* barrier implementation */
Ruleset p:Proc Do
  Rule "Advance release barrier"
    releaseflag[p] & L1_isRequestBufferEmpty(p)
    ==>
  Begin
    advanceRelease(p);
  Endrule;
Endruleset;

```

(a) Mur ϕ barrier rule.

```

Procedure advanceRelease(pid:Proc);
Begin
  If !spinwaiting[pid] then
    /* not spinwaiting => executing this
       barrier code for first time */
    /* a sense reversing barrier */
    localsense[pid] := !localsense[pid];
    /* no need for a lock here as
       Mur $\phi$ 's procedures are atomic */
    barcount := barcount + 1;
    If barcount = ProcCount then
      barcount := 0;
      barrier := localsense[pid];
      releaseflag[pid] := false;
      /* self-invalidate and unset touched
         bits after release is done */
      self_invalidate(pid);
      unset_touched(pid);
      /* the last proc to set the barcount
         to 0 also unsets the AccessStatus */
      unsetAccessStatus();
    Else
      spinwaiting[pid] := true;
    Endif;
  Else
    If barrier = localsense[pid] then
      releaseflag[pid] := false;
      spinwaiting[pid] := false;
      /* self-invalidate and unset
         touched bits after release is done */
      self_invalidate(pid);
      unset_touched(pid);
    Endif;
  Endif;
EndProcedure;

```

(b) Mur ϕ procedure for implementing a sense-reversing barrier.

Fig. 5. Mur ϕ code for end-of-phase implementation for DeNovo as a sense-reversing barrier. (a) Rule that gets triggered when inside the barrier indicated by *releaseflag* and empty *L1* request buffer (no outstanding requests) and (b) implementation of the sense-reversing barrier, including calls to end-of-phase operations like *self-invalidation* instruction and unsetting of *touched* bits.

are similar to the last three invariants of the MESI protocol. The last three invariants of the DeNovo protocol are checks on the *touched* bit functionality.

- Only one modifiable cache copy.* There cannot be two modifiable *L1* cache copies in the system at the same time. This invariant checks that there are never two *L1* caches in *Registered* state for the same line at the same time.
- Data value consistency at L1.* This invariant has two parts: (1) If *L1* is in *Valid* state and *touched* bit is set (value is read in this phase) and *L2* is also in *Valid* state, then the data values should be the same at both *L1* and *L2*. (2) If *L1* is in *Valid* state and *touched* bit is set and some other *L1* is in *Registered* state,⁴ the data values should match.
- Data value consistency at L2.* This invariant checks that when *L2* is in *Valid* state and *dirty* bit is not set, *L2*'s data value should be the same as at memory.
- Touched bit on a write.* On a write, this invariant checks that no other cache has the *touched* bit set to true. This verifies that the *touched* bit is implemented correctly.
- Touched bit on a read.* Similar to the above, on a read, this invariant checks that the only cache lines that can have the *touched* bit set to true (for cores other than the requestor) are the ones in *Valid* state.
- Unsetting touched bits.* Finally, this invariant checks that all the *touched* bits are set to false at the end of the phase.

4. RESULTS

Through model checking, we found six bugs (including two deadlock scenarios) in MESI and three bugs in DeNovo.⁵ A deadlock occurs when all the entities in the system (all *L1*s and *L2*) stop making any forward progress. Mur ϕ checks for deadlock by default. In MESI, all the bugs except one of the deadlocks were caused by protocol races between *L1* writebacks and other cache events. The other deadlock was caused by incorrect handling of clean replacement at the *L2*.

Most of these bugs found in MESI involved subtle data races and took several days to track, debug, and fix. All three bugs found in DeNovo were simple to fix and showed mistakes in translating our internal high level specification into the implementation (i.e., their solutions were already present in our internal high level description of the protocol).

Each of the bugs found in MESI and DeNovo is described in detail next. In all these descriptions, we consider a single address. $L1_{P1}$, $L1_{P2}$, and *L2* indicate the cache lines corresponding to this address in core *P1*, core *P2*, and *L2*, respectively. As mentioned in Section 2, we assume an in-cache directory at *L2*, and hence we use the words *directory* and *L2* interchangeably.

4.1. MESI Bugs

We first discuss the six bugs found in the MESI protocol. We list them in decreasing order of complexity and the amount of change to the code required to fix them.⁶

Bug 1. The first bug is caused by a race between an *L1* writeback and a write request by some other *L1*. Figure 6 shows the events that lead to this bug. Let us assume that initially $L1_{P1}$ is in *Modified* state, $L1_{P2}$ is in *Invalid* state, and *L2* records that the

⁴This is possible in DeNovo because registration at the other *L1* may have happened in a previous parallel phase.

⁵We note that these bugs were discovered in 2011 when DeNovo was (and still is) much less mature than the GEMS MESI protocol. The latter was released in 2007 and had been used by many researchers over the course of four years when we found the bugs.

⁶We confirmed the fixes for the MESI bugs in a personal email exchange with Arkaprava Basu, one of the developers of the GEMS simulation suite.

Fig. 6. MESI Bug 1 showing a race between an $L1$ writeback and a remote write request.

cache entry is modified in $L1_{P1}$. Then $L1_{P1}$ issues a replacement (event 1 in Figure 6) triggering a writeback (*PUTX*) and transitions to a transient state waiting for an acknowledgment to this writeback request. Meanwhile, $L1_{P2}$ issues a write request (event 2) triggering *GETX* to $L2$. $L2$ first receives *GETX* from $L1_{P2}$ (event 3). It forwards the request to $L1_{P1}$ and waits for an acknowledgment from $L1_{P2}$. $L1_{P1}$, on receiving the *GETX* request (event 4), forwards the data to $L1_{P2}$ and transitions to *Invalid* state. Then $L1_{P2}$, on receiving the data from $L1_{P1}$ (event 5) transitions to *Modified* state and unblocks the directory, which in turn records that the cache entry is now modified in $L1_{P2}$. But the writeback (*PUTX*) sent by $L1_{P1}$ is still in the network and it can reach the directory at any time because we have an unordered network (event 7), causing an error. For example, suppose $L1_{P1}$ later services a write request invalidating $L1_{P2}$ and the directory is appropriately updated (not shown in the figure). $L1_{P1}$'s writeback (*PUTX*) then reaches the directory, which is clearly an error. The bug was found when the writeback acknowledgment from $L2$ reached $L1_{P1}$, triggering a “missing transition” failure ($L1_{P1}$ does not expect a writeback acknowledgment in *Modified* state).

We solved this problem by not transitioning $L1_{P1}$ to *Invalid* state on receiving $L1_{P2}$'s *GETX* request. $L1_{P1}$ now sends *DATA* to $L1_{P2}$ as before, but continues to stay in the transient state, *M_I*. The write request from $L1_{P1}$, which triggered the bug in the previous example, is now kept pending because $L1_{P1}$ is in a transient state. We also added a transition at the $L2$ to send a writeback acknowledgment when the requester is not the owner in the directory's record. $L1_{P1}$ transitions to *Invalid* state on receiving the writeback acknowledgment from $L2$. With this, there is no longer a dangling *PUTX* in the network and the problem is solved. The trace for this bug involved multiple writes to the same memory location in a parallel phase. This scenario does not arise in DeNovo because the software guarantees data-race-freedom.

Bug 2. The second bug is similar to the first except that it is caused by a race between an $L1$ writeback and a read request by some other $L1$.

The first two bugs were the most complex to understand and fix. Most of the time was spent in discovering the root cause of the bugs and developing a solution in an

already complex protocol. The solutions to these bugs required adding two new cache events and eight new transitions to the protocol.

Bug 3. The third bug is caused by an unhandled protocol race between *L2* and *L1* replacements. To begin with, $L1_{P1}$ is in *Exclusive* state and *L2* records that *P1* is the exclusive owner. Then, both *L2* and *L1* replace the lines simultaneously, triggering invalidation and writeback messages, respectively. $L1_{P1}$, on receiving the invalidation message, transitions to *Invalid* state and sends its data to *L2*. On receiving this data, *L2* completes the rest of the steps for the replacement. In the end, both *L1* and *L2* have transitioned to *Invalid* states, but the initial writeback message from *L1* is still in the network and this is incorrect. The bug was found when the writeback acknowledgment (issued by *L2* on receiving the dangling writeback message) reaches $L1_{P1}$ when it is not expecting one and hence triggers a “missing transition” error.

This bug can be fixed by not sending the data when *L1* receives an invalidation message and by treating the invalidation message itself as the acknowledgment for *L1*’s earlier writeback message. Also, the *L1* writeback message is treated as the data response for the invalidation message at *L2*. The fix required adding four new transitions to the protocol.

Bug 4. The fourth bug results in a deadlock situation. It is caused by an incorrectly handled protocol race between an *Exclusive_unblock* (response sent to unblock *L2* on receiving an exclusive access) and an *L1* writeback issued by the same *L1* (issued after sending *Exclusive_unblock*). Initially, *L2* is waiting for an *Exclusive_unblock* in a transient state transitioned from *Invalid* state. In this transient state, when *L2* receives an *L1* writeback, it checks whether this writeback came from the current owner or not. The owner information is updated at *L2* on receiving the *Exclusive_unblock* message. Here, *L1* writeback (racing with *Exclusive_unblock* from the same *L1*) reached *L2* first and *L2* incorrectly discarded the *L1* writeback because the owner information at *L2* did not match the sender of the *L1* writeback. This incorrect discarding of the *L1* writeback results in a deadlock.

This bug can be fixed by holding the *L1* writeback to be serviced until *Exclusive_unblock* is received by *L2*. This requires adding a new transition and additional buffering to hold the stalled request to the protocol.

Bug 5. The fifth bug is similar to the fourth (race between *Exclusive_unblock* and *L1* writeback), but instead *L2* is initially in *Shared* state. The fix for this bug required adding two new transitions and additional buffering to hold the stalled requests to the protocol.

Bug 6. The last bug results in a deadlock scenario due to an incorrect transition by *L2* on a clean replacement. It transitions to a transient state awaiting an acknowledgment from memory even though the transition did not trigger any writeback. The fix was simple and required transitioning to *Invalid* state instead.

4.2. DeNovo Bugs

We next discuss the three bugs found in the DeNovo protocol. The first bug is a performance bug and the last two are correctness bugs, both of which are caused by races related to writebacks.

Bug 1. The first of the three bugs found was caused by not unsetting the dirty bit on replacement of a dirty *L2* cache line. Assume that *L2* is initially in *Valid* state and the *dirty* bit is set to true. Then, on *L2* replacement, it transitions to *Invalid* state and writes back data to memory. But the *dirty* bit is mistakenly not unset. This bug was found when Mur ϕ tried to replace the line in *Invalid* state as the *dirty* bit was

set to true (the model triggers a replacement event by only checking the *dirty* bit). The model, legitimately, did not have an action specified for a replacement event in the *Invalid* state, thus resulting in a “missing transition” error. However, the actual implementation did have an action (incorrectly) that triggered unnecessary writebacks to memory, which should be silent replacements instead. This turned out to be a rare case to hit in the simulation runs.

Bug 2. This occurs because an *L2* initiated writeback and future requests to the same cache line are not serialized. Initially, $L1_{P1}$ is in *Registered* state and *L2* knows $P1$ as the registrant. On replacing the line, *L2* sends a writeback request to *L1*. *L1* replies to this writeback request by sending the data to *L2* and transitions to *Valid* state.⁷ Then, on receiving the writeback from *L1*, *L2* sends an acknowledgment to *L1* and, in parallel, sends a writeback to memory and waits for an acknowledgment. Meanwhile, let us assume that *L1* issued a registration request (on receiving a store request) and successfully registers itself with *L2*. At this point, yet another *L2* replacement was triggered, finally leading to multiple writebacks to memory in flight. This is incorrect because the writebacks can be serviced out of order. Mur ϕ found this bug when an assertion failed inside the implementation of *L2*’s request buffer.

The real source of this bug is allowing *L1* registration to be serviced at *L2* while a writeback to memory is pending. The fix involves serializing requests to the same location at *L2*—in this case, the *L1* registration request behind the writeback to memory. This was already present in our high level specification but was missed in the actual protocol implementation. It did not involve adding any new states or transitions to the protocol.

Bug 3. The last bug is due to a protocol race where both the *L1*s and the *L2* replace the line. This bug involves both cores and cross-phase interactions. The events that lead to the bug are shown in Figure 7. At the beginning of the phase, let us assume that $L1_{P1}$ is in *Invalid* state and $L1_{P2}$ is in *Registered* state (from the previous phase). $L1_{P2}$ replaces the line (event 1 in Figure 7) and issues a writeback (PUTX) to *L2*. While this writeback is in flight, $L1_{P1}$ successfully registers itself with *L2* (events 2–4) (*L2* redirects the request to $L1_{P2}$ because it is the current registrant). This is followed by a replacement by $L1_{P1}$ (event 5), thus triggering another writeback (PUTX) to *L2*. *L2* first receives the writeback from $L1_{P1}$ (event 6) and responds by sending an acknowledgment and transitioning to *Valid* state while setting the *dirty* bit to true. Now, *L2* also replaces the line (event 7) transitioning to *Invalid* state and issues a writeback to memory. But the writeback from $L1_{P2}$ is still in flight. This writeback now reaches *L2* (event 8) while in *Invalid* state (because we model an unordered network). The implementation did not handle this case and resulted in a “missing transition” failure. This bug can be easily fixed by adding a transition to send an acknowledgment to $L1_{P2}$ ’s writeback without the need for triggering any actions at *L2*.

4.3. Analysis

The bugs just described for both MESI and DeNovo show that cache line replacements and writebacks, when interacting with other cache events, cause subtle races and add to the complexity of cache coherence protocols. Fixes to bugs in the MESI protocol needed adding new events and several new transitions. On the other hand, fixing bugs in the DeNovo protocol was relatively easy since it lacks transient states even for races related to writebacks.

⁷In DeNovo, the *L2* cache is inclusive of only the registered lines in any *L1*. Hence, it is possible for *L1* to transition from *Registered* to *Valid* on receiving a writeback request from *L2*.

5.1. Verification Techniques

Hardware coherence protocols have numerous transient states and hard-to-cover race conditions making it very difficult to find all the bugs using just simulations or random testing. Hence, formal methods like model checking are often employed to verify their correctness. Model checking is a technique to verify the properties of a system by exhaustive exploration of the state space [Clarke and Emerson 1982; Queille and Sifakis 1982]. McMillan and Schwalbe's seminal work on model checking the Encore Gigamax protocol [McMillan and Schwalbe 1991] was the first to apply model checking to verify cache coherence protocols.

Complex systems often exhibit a lot of regularity and symmetry. Ip and Dill developed Mur ϕ [Dill et al. 1992; Ip and Dill 1993, 1996], which exploits these characteristics by grouping together similar states to verify a reduced state graph instead of the full one. This helps to greatly reduce the amount of time and memory used in verification. Mur ϕ is a widely used tool to formally verify cache coherence protocols; for example, Sun RMO memory model [Park and Dill 1995], Sun S3.mp multiprocessor [Pong et al. 1998], Cray SV2 protocol [Abts et al. 2000], and Token coherence protocol [Burckhardt et al. 2005].

In this article, we use Mur ϕ for our protocol verification work. Using model checking tools like Mur ϕ for verifying cache coherence protocols is not new per se. This article, nevertheless, does discuss some extensions to the canonical coherence protocol modeling technique to model the guarantees provided by disciplined programming languages (namely, data-race-freedom and a disciplined parallel phase behavior). The main contribution of this work is to provide a detailed experience of the verification process of a state-of-the-art, publicly available, mature, modern hardware coherence protocol implementation (MESI). We further compare this experience with that of verifying a protocol driven by a hardware-software co-design approach (DeNovo), motivated by simplifying both software and hardware through the same mechanisms (while providing opportunities for performance and energy improvements).

Explicit state exploration model checking tools traditionally have the problem of state space explosion limiting the scalability of such tools. As mentioned in Section 4.4, we ran out of system memory when we increased the verification parameters. We tried a distributed model checker based on Mur ϕ , Preach [Bingham et al. 2010] (similar to Eddy Murphi [Melatti et al. 2009]). But PReach did not help us reduce the number of states explored, and hence the memory footprint stayed the same.

There are other verification techniques that do not have this state explosion problem and can scale to larger systems. Parametric verification [Chou et al. 2004; McMillan 2001; O'Leary et al. 2009] and theorem proving [Seungjoon and David 1996] are two such techniques. Several of these techniques are sometimes combined together to verify a given system. For example, the verification of Token Coherence [Burckhardt et al. 2005] is achieved by combining assume-guarantee reasoning and structural induction (in addition to model checking). Even though the techniques employed were able to verify the protocol, in practice, these techniques are either hard for nonspecialists to use or error-prone because of laborious manual intervention [Zhang et al. 2014]. There have been proposals to automate parametric verification techniques and minimize manual intervention [Chou et al. 2004], but such techniques impose severe limitations on the protocols that can be verified. DeNovo, in contrast, is a simpler protocol and makes it feasible to verify with easy-to-use verification techniques, such as an explicit state model checker. A general survey of various techniques used to verify cache coherence protocols can be found in Pong and Dubois [1997].

5.2. Redesigning Hardware Coherence Protocols

There has been some work on redesigning the hardware cache coherence protocol itself to address the verification problem. A recent proposal, PVCohereence [Zhang

et al. 2014], lists various guidelines for designing cache coherence protocols so that they can be verified using existing automatic parametric verification techniques. The verification technique used in the paper, Simple-PV, uses an automatic tool to generate the parametric model and then Mur ϕ to verify the generated model. When the proposed guidelines were applied to the MOESI protocol to make it amenable to parametric verification (in the process also making it even more complex), the authors noticed that the resulting protocol could not be verified by Mur ϕ and needed even more changes. The final protocol was verifiable but showed performance degradation (e.g., average 5% and up to 13.8% increase in network traffic). In Fractal Coherence [Zhang et al. 2010; Voskuilen and Vijaykumar 2014], the verification for arbitrary number of cores is made possible by just verifying the minimum system for correctness and verifying that the whole system has fractal behavior. But Fractal Coherence requires addition of states and messages to an existing cache coherence protocol (MOSI, as described in Zhang et al. [2010]) to maintain fractal behavior. Thus, it enables verification of the entire system at the expense of increasing the complexity of the base protocol. Also, a specific implementation of Fractal Coherence, TreeFractal, shows a performance degradation (>10%) when compared to traditional protocols. FlatFractal [Voskuilen and Vijaykumar 2014] proposes techniques to eliminate some of the limitations of TreeFractal, making FlatFractal's performance comparable to the base implementation on 16 cores. But when scaled to 32 cores, FlatFractal shows up to 40% degradation in total execution time. Atomic Coherence [Vantrease et al. 2011] attempts to simplify the coherence protocol by separating out the races from the protocol. The design avoids the protocol races by requiring each coherence action to be guarded by a mutex. It uses nanophotonics for performing these mutex operations with low latency because they are now on the critical path. As a result, Atomic Coherence eliminates transient states in the coherence protocol but heavily relies on a specific type of on-chip network.

In contrast to these approaches, DeNovo reduces the complexity of the cache coherence protocol by following a hardware-software co-design approach that eliminates races and thereby reduces the overall verification overhead. The basic DeNovo protocol is shown to be performance-competitive with MESI, and its simplicity enables optimizations that significantly improve performance and network traffic, thereby also potentially saving energy. The DeNovo coherence protocol used in this article assumes deterministic software. DeNovo is an ongoing research project seeking to broaden the class of software supported by such simple protocols. Recently, we extended DeNovo to support disciplined lock-based programs using dynamically generated access signatures to guide self-invalidations [Bocchino et al. 2011; Sung et al. 2013, 2014]. These extensions are rather small, and we believe can be verified independently. We are also developing simple extensions to support more complex lock-free synchronization algorithms. With these extensions, DeNovo will be able to support all codes correctly as long as synchronization operations are explicit.⁸ For legacy codes that do not contain annotations for self-invalidations or generating dynamic access signatures, performance (but not correctness) may be affected. DeNovo's vision of using disciplined parallel programming practices for better performance and programmability is shared by other projects as well, including for complex system code such as operating systems (e.g., Barrelfish [Baumann et al. 2009]). This paper provides an objective study to quantify the complexity advantages of such approaches.

6. CONCLUSION

With the era of manycores, there is a renewed debate in the community as to whether we have tamed the complexity of hardware coherence enough to employ it in systems

⁸Recent standards efforts for popular languages like Java, C, and C++ already require making synchronizations explicit.

with hundreds of cores or whether we should abandon it for other software alternatives. Although some have argued that hardware coherence is here to stay [Martin et al. 2012] and continue to work on even more complex extensions [Aisopos and Peh 2011], others are proposing systems that use software-directed coherence [Borkar 2011; Choi et al. 2011; Howard et al. 2010; IntelSCC 2009; Kelm et al. 2009]. This article adds objective results from a concrete case study to this debate. We attempted to formally verify a publicly available, state-of-the-art implementation of the MESI protocol. Surprisingly, we found that after four years of extensive use in the architecture community, the protocol implementation still had several bugs. These bugs were hard to diagnose and fix, requiring new state transitions. In contrast, verifying a far less mature, hardware-software co-designed protocol, DeNovo, revealed fewer bugs that were much easier to fix. After the bug fixes, we found that MESI took 20X longer to model check and required significant simplifying assumptions for the verification to be tractable. We conclude that even with a multitude of prior work, hardware coherence protocols are still difficult to design and understand. Furthermore, although it is difficult to define a single metric to quantify the relative complexity of protocols or to generalize from two design points, our results indicate that hardware-software co-designed protocols may offer a simpler alternative. This article also reinforces the necessity of employing formal verification techniques to prove the correctness of these protocols, but these techniques still require simplifying assumptions.

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