

# A SAT-Based Counterexample Guided Method for Unbounded Synthesis

**Abstract.** Reactive synthesis techniques based on constructing the winning region of the system have been shown to work well in many cases but suffer from state explosion in others. A different approach, proposed recently, applies SAT solvers in a counterexample guided framework to solve the synthesis problem. However, this method is limited to synthesising systems that execute for a bounded number of steps and is incomplete for synthesis with unbounded safety and reachability objectives. We present an extension of this technique to unbounded synthesis. Our method applies Craig interpolation to abstract game trees produced by counterexample-guided search in order to construct a monotonic sequence of may-losing regions. Experimental results based on SYNTCOMP 2015 competition benchmarks show this to be a promising alternative that solves some previously intractable instances.

## 1 Introduction

Reactive systems are ubiquitous in real-world problems such as circuit design, industrial automation, or device drivers. Automatic synthesis can provide a *correct by construction* controller for a reactive system from a specification. Reactive synthesis is formalised as a game between the *controller* and its *environment*. In this work we focus on safety games, in which the controller must prevent the environment from forcing the game into an error state.

The reactive synthesis problem is 2EXPTIME-complete so naive algorithms are infeasible on anything but simple systems. There are several techniques that aim to mitigate this complexity by representing states and transitions of the system symbolically [17, 3, 15]. These techniques incrementally construct a symbolic representation of the set of discovered winning states of the game. The downside of this approach is that keeping track of all discovered winning states can lead to a space explosion even when using efficient symbolic representation such as BDDs or CNFs.

An alternative approach, proposed by Narodytska et al. [16], is to eschew states and focus on *runs* of the game. The method works by exploring a subset of the concrete runs of the game and proving that these runs can be generalised into a winning strategy on behalf of one of the players. In contrast to other existing synthesis methods, it does not store, in either symbolic or explicit form, the set of winning states. Instead, it uses counterexample-guided backtracking search to identify a small subset of runs that are sufficient to solve the game.

This method has been shown to outperform BDDs on certain classes of games; however it suffers from an important limitation: it is only able to solve games

with a bounded number of rounds. In case of safety games, this means proving that the controller can keep the game within the safe region for a bounded number of steps. This is insufficient in most practical situations that require unbounded realizability.

In this paper, we extend the method by Narodytska et al. [16] to unbounded safety games. To this end we enhance the method with computational learning: every time the search algorithm discovers a new counterexample, the learning procedure analyzes the counterexample, extracting a subset of states winning for one of the players from it. The learning procedure ensures that reaching a fixed point in these sets is sufficient to establish unbounded realizability. Our method can be seen as a hybrid between the counterexample-guided algorithm by Narodytska et al. and methods based on winning set computation: we use the former to guide the search, while we rely on the latter to ensure convergence.

We evaluate our method on the benchmarks of the 2015 synthesis competition (SYNTCOMP’15). While our solver solves fewer total instances than competitors, it solves the largest number of unique instances, i.e., instances that could not be solved by any other solver. These results confirm that there exist classes of problems that are hard for traditional synthesis techniques, but can be efficiently solved by our method. While further performance improvements are clearly needed, in its current state the method can be used in a portfolio solver to increase the number of tractable instances.

Section 2 outlines the original bounded synthesis algorithm. In Section 3 we describe and prove the correctness of our extension of the algorithm to unbounded games. In the following sections we discuss optimisations to the algorithm, evaluate our methodology, and compare our approach to other synthesis techniques.

## 2 Background

A *safety game*,  $G = \langle X, U, C, \delta, I, E \rangle$ , is defined over boolean state variables  $X$ , uncontrollable action variables  $U$ , and controllable action variables  $C$ .  $I$  is the initial state of the game given as a valuation of state variables.  $E(X)$  is the set of error states represented by its characteristic formula. The transition relation  $\delta(X, U, C, X')$  of the game is a boolean formula that relates current state and action to the set of possible next states of the game. We assume deterministic games, where  $\delta(x, u, c, x'_1) \wedge \delta(x, u, c, x'_2) \implies (x'_1 = x'_2)$ .

At every round of the game, the *environment* picks an uncontrollable action, the *controller* responds by choosing a controllable action and the game transitions into a new state according to  $\delta$ . A *run* of a game  $(x_0, u_0, c_0), (x_1, u_1, c_1) \dots (x_n, u_n, c_n)$  is a chain of state and action pairs s.t.  $\delta(x_k, u_k, c_k, x_{k+1})$ . A run is winning for the controller if  $x_0 = I \wedge \forall i \in \{1..n\}(\neg E(x_i))$ . In a *bounded game* with bound  $\kappa$  all runs are restricted to length  $\kappa$ , whereas unbounded games consider runs of infinite length. Since we consider only deterministic games, a run is uniquely described by a list of assignments to  $U$  and  $C$ .

A *controller strategy*  $\pi^c : (X, U) \rightarrow C$  is a mapping of states and uncontrollable inputs to controllable actions. A controller strategy is winning in a bounded game of rank  $n$  if all runs  $(x_0, u_0, \pi^c(x_0, u_0)), (x_1, u_1, \pi^c(x_1, u_1)) \dots (x_n, u_n, \pi^c(x_n, u_n))$  are winning. Bounded *realizability* is the problem of determining the existence of such a strategy for a bounded game.

An *environment strategy*  $\pi^e : X \rightarrow U$  is a mapping of states to uncontrollable actions. A bounded run is winning for the environment if  $x_0 = I \wedge \exists i \in \{1..n\}(E(x_i))$  and an environment strategy is winning for a bounded game if all runs  $(x_0, \pi^e(x_1), c_1), (x_1, \pi^e(x_1), c_1) \dots (x_n, \pi^e(x_n), c_n)$  are winning for the environment. Safety games are zero sum, therefore the existence of a winning controller strategy implies the nonexistence of a winning environment strategy and vice versa.

## 2.1 Counterexample-guided bounded synthesis

We review the bounded synthesis algorithm by Narodytska et al. [16], which is the main building block for our unbounded algorithm.

Realisability for a bounded safety game can be solved by checking a quantified formula consisting of the game's transition relation *unrolled* to the length of the bounded game. The quantifiers represent the two players of the game: universal for the controller and existential for the environment. Just as the players alternate in choosing actions every step of the game, the formula is prefixed by a quantifier alternation over  $U$  and  $C$  for each step.

Instead of solving both quantifiers at once, two competing solvers are used to solve abstractions of the game. Each solver builds a candidate strategy for its corresponding player and uses its opponent's candidate to restrict its own search. The solvers communicate to refine one another's candidate strategies by finding counterexamples. Eventually one solver will be unable to find a counterexample, indicating that its opponent's candidate strategy is a winning strategy for the game.

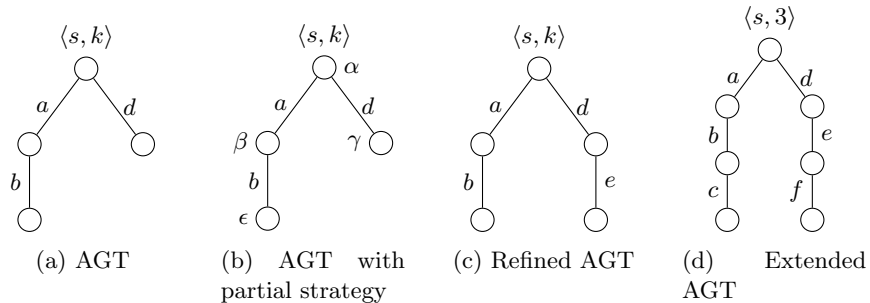


Fig. 1: Abstract game trees

An abstraction of the game restricts actions available to one of the players. Specifically, we consider abstractions represented as trees of actions, referred to as *abstract game trees*. Figure 1a shows an example abstract game tree for the controller (abstract game trees for the environment are similar). In the abstract game, the controller is required to pick actions from the tree, starting from the root node. After reaching a leaf, it continues playing unrestricted. The tree in Figure 1a restricts the initial controller action to the set  $\{a, d\}$ . After choosing action  $a$ , the controller is required to play action  $b$  in the next round. It then reaches a leaf of the tree and continues playing unrestricted. The root of the tree is annotated by the initial state  $s$  of the abstract game and the bound  $k$  on the number of rounds. We denote  $\text{NODES}(T)$  the set of all nodes of a tree  $T$  and  $\text{LEAVES}(T)$  the subset of leaf nodes.

Given a controller (environment) abstract game tree  $T$  a *partial strategy*  $\text{Strat} : \text{NODES}(T) \rightarrow U$  ( $\text{Strat} : \text{NODES}(T) \rightarrow C$ ) labels each node of the tree with the opponent’s action to be played in that node. Figure 1b shows an example partial strategy. The environment starts by choosing action  $\alpha$ . If the controller plays  $a$ , the environment responds with  $\beta$  in the next round, and so on. Given a partial strategy  $\text{Strat}$ , we can map each leaf  $l$  of the abstract game tree to  $\langle s', i' \rangle = \text{OUTCOME}(\langle s, i \rangle, \text{Strat}, l)$  obtained by playing all controllable and uncontrollable actions on the path from the root to the leaf.

The bounded synthesis algorithm solves abstract game trees by invoking a SAT solver to search for candidate partial strategies. Candidate strategies are checked and by a call to the opponent solver, which also refines the abstract game tree. The refinements are then solved recursively. The full procedure is illustrated in Algorithm 1.

The algorithm takes a concrete game  $G$  with maximum bound  $\kappa$  as an implicit argument. In addition, it takes a player  $p$  (controller or environment), state  $s$ , bound  $k$  and an abstract game tree  $T$  and returns a winning partial strategy for  $\text{OPPONENT}(p)$ , if one exists. The initial invocation of the algorithm takes the initial state  $I$ , bound  $\kappa$  and an empty abstract game tree  $\emptyset$ . Initially the solver is playing on behalf of the environment since that player takes the first move in every step. The empty game tree does not constrain opponent moves, hence solving such an abstraction is equivalent to solving the original concrete game.

The algorithm is organised as a counterexample-guided abstraction refinement (CEGAR) loop. The first step of the algorithm uses the `FINDCANDIDATE` function, described below, to come up with a candidate partial strategy for  $T$ . If it fails to find a strategy, this means that no winning partial strategy exists for  $T$ . If, on the other hand, a candidate partial strategy is found, we need to verify if it is indeed winning for  $T$ .

The `VERIFY` procedure searches for a *spoiling* counterexample strategy in each leaf of the candidate partial strategy by calling `SOLVEABSTRACT` for the opponent. The dual solver solves games on behalf of the controller player.

If the dual solver can find no spoiling strategy at any of the leaves, then the candidate partial strategy is a winning one. Otherwise, `VERIFY` returns the move used by the opponent to defeat a leaf of the partial strategy, which is appended

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**Algorithm 1** Bounded Synthesis

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1: function SOLVEABSTRACT( $p, s, k, T$ )
2:    $cand \leftarrow \text{FINDCANDIDATE}(p, s, k, T)$  ▷ Look for a candidate
3:   if  $k = 1$  then return  $cand$  ▷ Reached the bound
4:    $T' \leftarrow T$ 
5:   loop
6:     if  $cand = \emptyset$  then return  $\emptyset$  ▷ No candidate: return with no solution
7:      $\langle cex, l, u \rangle \leftarrow \text{VERIFY}(p, s, k, T, cand)$  ▷ Verify candidate
8:     if  $cex = \text{false}$  then return  $cand$  ▷ No counterexample: return candidate
9:      $T' \leftarrow \text{APPEND}(T', l, u)$  ▷ Refine AGT with counterexample
10:     $cand \leftarrow \text{SOLVEABSTRACT}(p, s, k, T')$  ▷ Solve refined AGT
11:   end loop
12: end function

13: function FINDCANDIDATE( $p, s, k, T$ )
14:    $\hat{T} \leftarrow \text{EXTEND}(T)$  ▷ Extend the tree with unfixed actions
15:    $f \leftarrow \text{if } p = \text{cont} \text{ then } \text{TREEFORMULA}(\hat{T}) \text{ else } \overline{\text{TREEFORMULA}(\hat{T})}$ 
16:    $sol \leftarrow \text{SAT}(s(X_{\hat{T}}) \wedge f)$ 
17:   if  $sol = \text{unsat}$  then
18:     if  $\text{unbounded}$  then ▷ Active only in the unbounded solver
19:        $\sigma \leftarrow \text{GENERALISE}(s)$  ▷ Expand  $s$  to a set of states
20:       if  $p = \text{cont}$  then  $\text{LEARN}(\sigma, \hat{T})$  else  $\overline{\text{LEARN}}(\sigma, \hat{T})$ 
21:     end if
22:     return  $\emptyset$  ▷ No candidate exists
23:   else
24:     return  $\{\langle n, c \rangle \mid n \in \text{NODES}(T), c = \text{SOL}(n)\}$  ▷ Fix candidate moves in  $T$ 
25:   end if
26: end function

27: function VERIFY( $p, s, k, T, cand$ )
28:   for  $l \in \text{leaves}(gt)$  do
29:      $\langle k', s' \rangle \leftarrow \text{OUTCOME}(s, k, cand, l)$  ▷ Get rank and state at leaf
30:      $u \leftarrow \text{SOLVEABSTRACT}(\text{OPPONENT}(p), k', s', \emptyset)$  ▷ Solve for the opponent
31:     if  $u \neq \emptyset$  then return  $\langle \text{false}, l, u \rangle$  ▷ Return counterexample
32:   end for
33:   return  $\langle \text{true}, \emptyset, \emptyset \rangle$ 
34: end function

```

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to the corresponding node in  $T$  in order to refine it in line (9) as shown in Figure (1c).

We solve the refined game by recursively invoking `SOLVEABSTRACT` on it. If no partial winning strategy is found for the refined game then there is also no partial winning strategy for the original abstract game, and the algorithm returns a failure. Otherwise, the partial strategy for the refined game is *projected* on the original abstract game by removing the leaves introduced by refinements. The resulting partial strategy becomes a candidate strategy to be verified at the next iteration of the loop. In the worst case the loop terminates after all actions in the game are refined into the abstract game.

The CEGAR loop depends on the ability to guess candidate partial strategies in `FINDCANDIDATE`. For this purpose we use the heuristic that a partial strategy may be winning if each `OUTCOME` of the strategy can be extended to a run of the game that is winning for the current player. Clearly, if such a partial strategy does not exist then no winning partial strategy can exist for the abstract game tree. We can formulate this heuristic as a SAT query, which is constructed recursively by `TREEFORMULA` (for the controller) or `TREEFORMULA` (for the environment) in Algorithm 2.

The tree is first extended to the maximum bound with edges that are labeled with arbitrary opponent actions (Algorithm 1, line 14) as shown in Figure ???. Then copies of state and action variables are introduced for each node in tree and opponent action variables for each edge. Opponent actions are set to `ACTION( $e$ )`, the action that labels the edge of the tree  $e$ . The new state and action variables are connected to the successor nodes (`SUCC`) by an encoding transition relation and constrained to the winning condition of the player.

### 3 Unbounded Synthesis

Bounded synthesis can be used to prove the existence of a winning strategy for the environment on the unbounded game by providing a witness. For the controller, the strongest claim that can be made is that the strategy is winning as long as the game does not extend beyond the maximum bound.

It is possible to set a maximum bound such that all runs in the unbounded game will be considered. The naïve approach is to use size of the state space as the bound ( $2^{|X|}$ ) so that all states may be explored by the algorithm. A more nuanced approach is to use the diameter of the game [2], which is the smallest number  $d$  such that for any state  $x$  there is a path of length  $\leq d$  to all other reachable states. However, the diameter is difficult to estimate and can still lead to infeasibly long games.

We instead present an approach that iteratively solves games of increasing bound while learning states. We show that reaching a fixed point in learned states is sufficient for completeness.

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**Algorithm 2** Tree formulas for Controller and Environment respectively
 

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1: function TREEFORMULA( $T$ )
2:   if BOUND( $T$ ) = 0 then
3:     return  $\neg E(X_T)$ 
4:   else
5:     return  $\neg E(X_T) \wedge$ 
6:       
$$\bigwedge_{\langle e, n \rangle \in \text{succ}(T)} (\delta(X_T, U_e, C_n, X_n) \wedge U_e = \text{ACTION}(e) \wedge \text{TREEFORMULA}(n))$$

7:   end if
8: end function
9: function TREEFORMULA( $T$ )
10:  if BOUND( $T$ ) = 0 then
11:    return  $E(X_T)$ 
12:  else
13:    return  $E(X_T) \vee$ 
14:      
$$\bigvee_{\langle e, n \rangle \in \text{succ}(T)} (\delta(X_T, U_n, C_e, X_n) \wedge C_e = \text{ACTION}(e) \wedge \overline{\text{TREEFORMULA}(n)})$$

15:  end if
16: end function

```

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### 3.1 Learning States with Interpolants

We modify the bounded synthesis algorithm to learn a subset of states from failed attempts to find candidate strategies.

When the abstract game is unsatisfiable this indicates that the states represented by nodes in the abstract game tree are losing for the current player. We extract these states from the game tree using interpolation.

Given two formulas  $F_1$  and  $F_2$  such that  $F_1 \wedge F_2$  is unsatisfiable, it is possible to construct a Craig interpolant [7]  $\mathcal{I}$  such that  $F_1 \rightarrow \mathcal{I}$ ,  $F_2 \wedge \mathcal{I}$  is unsatisfiable, and  $\mathcal{I}$  refers only to the intersection of variables in  $F_1$  and  $F_2$ . An interpolant can be constructed efficiently from a resolution proof of the unsatisfiability of  $F_1 \wedge F_2$  [18].

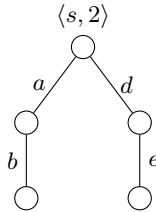


Fig. 2: A controller-losing game tree

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**Algorithm 3** Unbounded Synthesis

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1: function SOLVEUNBOUNDED( $T$ )
2:    $B^M \leftarrow E$ 
3:    $B^m[0] \leftarrow E$ 
4:   for  $k = 1 \dots$  do
5:     if  $\text{SAT}(I \wedge B^M)$  then return unrealisable  $\triangleright$  Losing in the initial state
6:     if  $\exists i \ B^m[i] \equiv B^m[i + 1]$  then return realisable  $\triangleright$  Reached fixed point
7:      $B^m[k] \leftarrow \top$ 
8:      $\text{CHECKRANK}(I, k)$ 
9:   end for
10: end function
Require: May and must invariants hold
Ensure: May and must invariants hold
Ensure:  $\neg s \in B^m[k]$  if there is a winning controller strategy of length  $k$  starting at  $s$ 
Ensure:  $s \in B^M$  if there is a winning environment strategy of length  $k$  starting at  $s$ 
11: function CHECKRANK( $s, k$ )
12:   return SOLVEABSTRACT( $\text{env}, \langle s, k \rangle, \emptyset$ )
13: end function

```

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Consider the game tree in Figure 2. The tree is losing for the controller, the topmost node is at rank 1, and the  $a$  and  $b$  nodes are at rank 0. Since the tree is controller-losing we know that at least one state in the tree will uncontrollably lead to the error state. As a result,  $\text{TREEFORMULA}(T)$  is unsatisfiable. If we take a step of the tree, such as the step where the environment plays  $a$ , and cut it from the rest of the tree then  $\text{TREEFORMULA}(\text{step}) \wedge \text{TREEFORMULA}(\text{parent})$  must too be unsatisfiable.

We can construct an interpolant with  $F_1 = \text{TREEFORMULA}(\text{parent})$  and  $F_2 = \text{TREEFORMULA}(\text{step})$ . The only variables shared between  $F_1$  and  $F_2$  are the state variables at rank 1. We know that  $F_2 \wedge \mathcal{I}$  is unsatisfiable, therefore all states in  $\mathcal{I}$  must lose to the uncontrollable action  $a$ . We also know that  $F_1 \rightarrow \mathcal{I}$ , thus  $\mathcal{I}$  contains all states reachable by the parent tree (on runs that avoid the error state.)

Now we can consider the step where the environment plays  $b$ , and the parent now without either step in the snippet. The formula  $(\text{TREEFORMULA}(\text{parent}) \wedge \mathcal{I}) \wedge \text{TREEFORMULA}(\text{step})$  must be unsatisfiable.  $\mathcal{I}$  contains all states that lose to  $a$  so any other state reachable at rank 1 in this subtree must lose to  $b$ . Therefore we can compute another interpolant that contains states that lose to  $b$ .

This is the foundation for a recursive algorithm that consumes an entire tree by removing a single step on each iteration (Algorithm 5). All learned states from which the controller must lose are recorded in a set of bad states  $B^M$ . Environment-losing are learned in the same way but are recorded by removing the state from a set of sets of controller may-losing states  $B^m$ . An environment-losing state of rank  $n$  is removed from every set at rank  $n$  and lower.



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**Algorithm 4** Amended tree formulas for Controller and Environment respectively

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1: function TREEFORMULA(gt)
2:   if RANK(gt) = 0 then
3:     return  $\neg B^M(x^{gt})$ 
4:   else
5:     return  $\neg B^M(x^{gt}) \wedge \bigwedge_{n \in \text{succ}(gt)} (\delta(n) \wedge \text{LABEL}(n) \wedge \text{TREEFORMULA}(n))$ 
6:   end if
7: end function
8: function  $\overline{\text{TREEFORMULA}}$ (gt)
9:   if RANK(gt) = 0 then
10:    return  $E(x^{gt})$ 
11:   else
12:    return  $B^m[\text{rank}(gt)](x^{gt}) \wedge$ 
13:     $(E(x^{gt}) \vee \bigvee_{n \in \text{succ}(gt)} (\delta(n) \wedge \text{LABEL}(n) \wedge \overline{\text{treeFormula}}(n)))$ 
14:   end if
15: end function

```

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We compile interpolants into DNF cubes by constructing a temporary BDD. This smaller representation is necessary for maintaining the size of formulas passed to the SAT solver. However, it can occasionally lead to state explosion.

### 3.2 Correctness

The correctness of the unbounded synthesis algorithm can be established independently from that of the bounded algorithm. Fortunately, the correctness of the bounded solver has been established in [16].

We define two global invariants of the algorithm. The *may-invariant* states that sets  $B^m[i]$  grows monotonically with  $i$  and that each  $B^m[i+1]$  overapproximates the uncontrollable predecessor of  $B^m[i]$ :

$$\forall i < k. B^m[i] \subseteq B^m[i+1], \text{Upred}(B^m[i]) \subseteq B^m[i+1].$$

The *must-invariant* guarantees that the must-losing set  $B^M$  is an underapproximation of the actual losing set  $B$ :

$$B^M \subseteq B.$$

The sets  $B^m$  and  $B^M$  are only modified by the inductive solver, implemented by LEARN and  $\overline{\text{LEARN}}$  functions. Below we prove that these functions indeed maintain the invariants.

### 3.3 Proof of $\overline{\text{learn}}$

We prove that postconditions of  $\overline{\text{LEARN}}$  are satisfied assuming that its preconditions hold.

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**Algorithm 5** Learning algorithms
 

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**Require:**  $\sigma(X_T) \wedge \text{TREEFORMULA}(T) \equiv \perp$ 
**Require:** *Must-invariant* holds

**Ensure:** *Must-invariant* holds

**Ensure:**  $\sigma(X_T) \wedge B^M$ 

 1: **function** LEARN( $\sigma, T$ )

 2:   **if**  $T = \emptyset$  **then return**

 3:    $n \leftarrow$  non-leaf node with max rank

 4:    $\langle T_1, T_2 \rangle \leftarrow \text{GTSPLIT}(T, n)$ 

 5:    $\mathcal{I} \leftarrow \text{INTERPOLATE}(\sigma(X_T) \wedge \text{TREEFORMULA}(T_1), \text{TREEFORMULA}(T_2))$ 

 6:    $B^M \leftarrow B^M \vee \mathcal{I}$ 

 7:   LEARN( $\sigma, T_1$ )

 8: **end function**
**Require:**  $\sigma(X_T) \wedge \overline{\text{TREEFORMULA}(T)} \equiv \perp$ 
**Require:** *May-invariant* holds

**Ensure:** *May-invariant* holds

**Ensure:**  $\sigma(X_T) \wedge B^m[\text{RANK}(T)] \equiv \perp$ 

 9: **function** LEARN( $\sigma, T$ )

 10:   **if**  $T = \emptyset$  **then return**

 11:    $n \leftarrow$  non-leaf node with max rank

 12:    $\langle T_1, T_2 \rangle \leftarrow \text{GTSPLIT}(T, n)$ 

 13:    $\mathcal{I} \leftarrow \text{INTERPOLATE}(\sigma(X_T) \wedge \overline{\text{TREEFORMULA}(T_1)}, \overline{\text{TREEFORMULA}(T_2)})$ 

 14:   **for**  $i = 1$  to  $\text{RANK}(n)$  **do**

 15:      $B^m[i] \leftarrow B^m[i] \wedge \neg \mathcal{I}$ 

 16:   **end for**

 17:    $\overline{\text{LEARN}}(\sigma, T_1)$ 

 18: **end function**


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Line (11) splits the tree  $T$  into  $T_1$  and  $T_2$ , such that  $T_2$  has depth 1. Consider formulas  $F_1 = \sigma(X_T) \wedge \overline{\text{TREEFORMULA}}(T_1)$  and  $F_2 = \overline{\text{TREEFORMULA}}(T_2)$ . These formulas only share variables  $X_n$ . Their conjunction  $F_1 \wedge F_2$  is unsatisfiable, as by construction any solution of  $F_1 \wedge F_2$  also satisfies  $\sigma(X_T) \wedge \overline{\text{TREEFORMULA}}(T)$ , which is unsatisfiable (precondition (b)). Hence the interpolation operation is defined for  $F_1$  and  $F_2$ .

Intuitively, the interpolant computed in line (13) overapproximates the set of states reachable from  $\sigma$  by following the tree from the root node to  $n$ , and underapproximates the set of states from which the environment loses against tree  $T_2$ .

Formally,  $\mathcal{I}$  has the property  $\mathcal{I} \wedge F_2 \equiv \perp$ . Since  $T_2$  is of depth 1, this means that the environment cannot force the game into  $B^m[\text{RANK}(n)]$  playing against the counterexample moves in  $T_2$ . Hence,  $\mathcal{I} \cap \text{Upred}(B^m[\text{RANK}(n)]) = \emptyset$ . Furthermore, since the may-invariant holds,  $\mathcal{I} \cap \text{Upred}(i) = \emptyset$ , for all  $i < \text{RANK}(n)$ . Hence, removing  $\mathcal{I}$  from all  $B^m[i]$ ,  $i \leq \text{RANK}(n)$  in line (15) preserves the may-invariant, thus satisfying the first post-condition.

Furthermore, the interpolant satisfies  $F_1 \rightarrow \mathcal{I}$ , i.e., any assignment to  $X_n$  that satisfies  $\sigma(X_T) \wedge \overline{\text{TREEFORMULA}}(T_1)$  also satisfies  $\mathcal{I}$ . Hence, removing  $\mathcal{I}$  from  $B^m[\text{RANK}(n)]$  makes  $\sigma(X_T) \wedge \overline{\text{TREEFORMULA}}(T_1)$  unsatisfiable, and hence all preconditions of the recursive invocation of  $\overline{\text{LEARN}}$  in line (17) are satisfied.

At the second last recursive call to  $\overline{\text{LEARN}}$ , tree  $T_1$  is empty,  $n$  is the root node,  $\overline{\text{TREEFORMULA}}(T_1) \equiv B^m[\text{RANK}(T)](X^T)$ ; hence  $\sigma(X_T) \wedge \overline{\text{TREEFORMULA}}(T_1) \equiv \sigma(X_T) \wedge B^m[\text{RANK}(T)](X^T) \equiv \perp$ . Thus the second postcondition of  $\overline{\text{LEARN}}$  holds.

### 3.4 Proof of Termination

We must prove that  $\text{CHECKRANK}$  terminates and that upon termination its postcondition holds.

We must prove that  $\text{CHECKRANK}$  terminates and that upon termination its postcondition holds, i.e., state  $s$  is removed from  $B^m[k]$  if there is a winning controller strategy on the bounded safety game of rank  $k$  or it is added to  $B^M$  otherwise. Termination follows from completeness of counterexample-guided search, which terminates after enumerating all possible opponent moves in the worst case.

Assume that there is a winning strategy for the controller at rank  $k$ . This means that at some point the algorithm discovers a counterexample tree of rank  $k$  for which the environment cannot force into  $E$ . The algorithm then invokes the  $\overline{\text{LEARN}}$  method, adds  $s$  to  $B^M$ . Alternatively, if there is a winning strategy for the environment at rank  $k$  then a counterexample losing for the controller will be found. Subsequently  $\text{LEARN}$  will be called and  $s$  eliminated from  $B^m[k]$ .

## 4 Optimisations

### 4.1 Generalising the initial state

If more states are removed from  $B^m$  on each iteration of  $\text{CHECKRANK}$  the algorithm will converge and terminate sooner. Additionally, any states removed

from  $B^M$  will reduce the number of states to be considered by the controller in future iterations.

As it is written, the algorithm only considers an overapproximations of states reachable from  $I$  for learning. Assuming that the algorithm does not terminate, then the addition of an extra step into the game on each iteration has the potential to greatly increase the reachable states. Considering some of these states earlier than they become reachable may lead to earlier termination.

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**Algorithm 6** Generalise  $I$  optimisation

---

```

function CHECKRANK( $s, k$ )
   $r \leftarrow \text{SOLVEABSTRACT}(\text{env}, \langle s, k \rangle, \emptyset)$ 
  if  $r \neq \emptyset$  then return  $r$ 
   $s' \leftarrow s$ 
  for  $x \in s$  do
     $r \leftarrow \text{SOLVEABSTRACT}(\text{env}, \langle s' \setminus \{x\}, k \rangle, \emptyset)$ 
    if  $r = \emptyset$  then  $s' \leftarrow s' \setminus \{x\}$ 
  end for
  return  $\emptyset$ 
end function

```

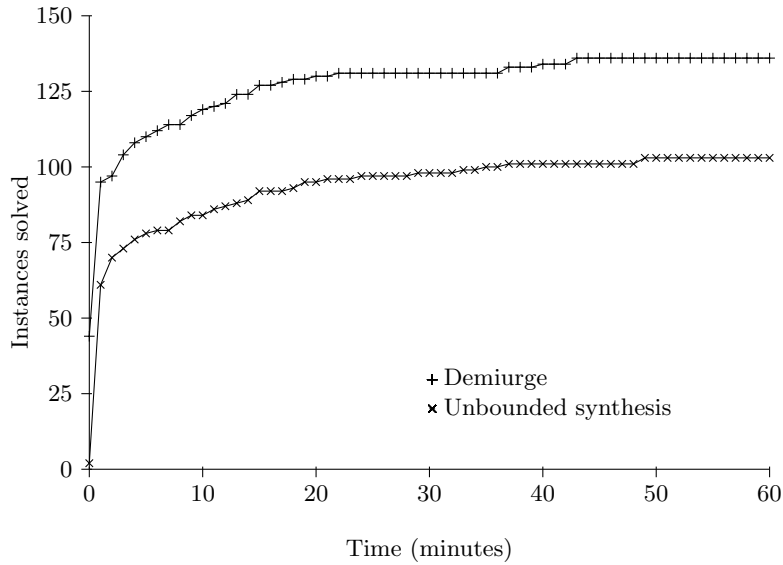
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The optimisation that allows this is relatively simple and is inspired by a common greedy heuristic for minimising **unsat** cores.  $I$  is a value assignment to each variable in  $X$ . If the environment does not win for  $\langle I, k \rangle$  then we attempt to solve for a generalised version of  $I$  by removing one assignment at a time. If the controller can win from the larger set of states then we continue generalising without it. In this way we learn more states by increasing the reachable set. In our benchmarks we have observed that this optimisation is beneficial on the first few iterations of CHECKRANK.

## 5 Evaluation

We evaluate our approach on the benchmarks of the 2015 synthesis competition (SYNTCOMP'15). Each benchmark comprises of controllable and uncontrollable inputs to a circuit that assigns values to latches. One latch is configured as the error bit that determines the winner of the safety game. The benchmark suite is a collection of both real-world and toy specifications including generalised buffers, AMBA bus controllers, device drivers, and converted LTL formulas. Descriptions of many of the benchmark families used can be found in the 2014 competition report [12].

The implementation of our algorithm uses GLUCOSE [1] for SAT solving and PERIPLO [19] for interpolant generation. We compared our implementation to DEMIURGE [3], the only SAT-based solver to compete in the sequential realisability track of SYNTCOMP'15. The benchmarks were run on a cluster of Intel



Quad Core Xeon E5405 CPUs with 16GB of memory. Each solver was allowed exclusive access to a node for one hour to solve an instance.

Our implementation was able to solve 103 out of the 250 specification in the allotted time, including 12 instances that were unable to be solved by any other solver in SYNTCOMP'15. DEMIURGE was able to solve 136 specifications. 77 of those instances were also solved by our implementation, so DEMIURGE solved 59 instances that our solver did not. However, our method was able to solve 26 specifications not solved by DEMIURGE.

For reference, the results of the synthesis competition are shown in Figure 3. The competition ran on better hardware with a longer time period, which explains why DEMIURGE performed better.

Solver	Solved	Unique
Simple BDD Solver (2)	195	10
AbsSynthe (seq2)	187	2
Simple BDD Solver (1)	185	
AbsSynthe (seq3)	179	
Realizer (sequential)	179	
AbsSynthe (seq1)	173	1
Demiurge (D1real)	139	5
Aisy	98	
Unbounded Synthesis	103	12

Fig. 3: Synthesis Competition 2015 Results

Five of the instances unique to our solver are device driver instances and another five are from the **stay** family, which are classified as toy instances. Out of the 26 instances that DEMIURGE could not solve, there were only 6 different benchmark classes. This supports the hypothesis that different game solving methodologies perform better on certain classes of specifications.

Whilst our solver is unable to solve as many instances as other tools yet, it was able to solve more unique instances than any solver in the competition. This shows that the method has potential. In the parallel track, DEMIURGE competes with a suite of 3 separate but communicating solvers. The solvers relay unsafe states to one another, which is compatible with the set  $B^M$  in our solver. It remains future work to explore the possibility of combining these solvers.

## 6 Related Work

Synthesis of safety games is a thoroughly explored area of research with most efforts directed toward solving games with BDDs [6] and abstract interpretation [20, 5]. Satisfiability solving has been used previously for synthesis in a suite of methods proposed by Bloem et al [3]. The authors propose a propose employing competing SAT solvers to learn clauses, which is similar to our approach but does not unroll the game. They also suggest QBF solver, template-based, and Effectively Propositional Logic (EPL) approaches.

SAT-based bounded model checking approaches that unroll the transition relation have been extended to unbounded by using conflicts in the solver [13], or by interpolation [14]. However, there are no corresponding adaptations to synthesis. Incremental induction [4] is another technique for unbounded model checking with an equivalent synthesis method [15], which computes the set of safe states incrementally.

There are different approaches to bounded synthesis than the one described here. The authors of [11] suggest a methodology directly inspired by bounded model checking and it has been adapted to symbolic synthesis [9]. Lazy synthesis [10] is a counterexample-guided approach to bounded synthesis that refines an implementation for the game instead of an abstraction of it.

The algorithm presented in this paper solves realisability for safety games. There is a method for extracting strategies from abstract game trees that is compatible with our method [8]. It involves a similar interpolation approach for discovering states in game tree nodes.

## 7 Conclusion

We presented an extension to an existing bounded synthesis technique that promotes it to unbounded safety games. The approach taken as whole differs from other synthesis techniques by combining counterexample guided game solving with winning set computation. The results show that our approach is able to solve more unique instances than other solvers by performing well on certain classes of games that are hard for other methodologies.

Future work includes incorporating the solver into a parallel suite of communicating solvers [3]. There is evidence that different solvers perform well on different classes of games. Thus we hypothesise that the way forward for synthesis tools is to combine the efforts of many different techniques.

We also aim to experiment with optimisations by focusing on heuristics for avoiding unhelpful actions as well as constructing succinct representations of interpolants without resorting to BDDs.

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