On Supervalid Inequalities for Binary Interdiction Games

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Abstract

Supervalid inequalities are a specific type of constraints often used within the branch-and-cut framework to strengthen the linear relaxation of mixed-integer programs. These inequalities share the particular characteristic of potentially removing feasible integer solutions as long as they are already dominated by an incumbent solution. This paper focuses on supervalid inequalities for solving binary interdiction games. Specifically, we provide a general characterization of inequalities that are derived from bipartitions of the leader's strategy set and develop an algorithmic approach to use them. This includes the design of two verification subroutines that we apply for separation purposes. We provide three general examples in which we apply our results to solve binary interdiction games targeting shortest paths, spanning trees, and vertex covers. Finally, we prove that the separation procedure is efficient for the class of interdiction games defined on greedoids — a type of set system that generalizes many others such as matroids and antimatroids.

Keywords: Network interdiction, Cutting planes, Integer programming, Greedoids.

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1 Introduction

This paper studies a broad class of Stackelberg competitions called binary interdiction games, where two players, denoted the *leader* and *follower*, sequentially solve interdependent optimization problems with conflicting objectives. Drawing from the definition and notation established in [59], binary interdiction games can be characterized as follows. Given a set of elements Δ representing the ground set of the game, let collections $\Pi \subseteq 2^{\Delta}$ and $\Omega \subseteq 2^{\Delta}$ denote the solution spaces of the leader and follower, respectively. Specifically, set Ω, called the *structure set*, comprises a collection

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of subsets of Δ satisfying some structural properties, from which the follower selects one that minimizes a weight function $w : \Delta \to \mathbb{R}^+$. In turn, set Π, denoted the *strategy set*, contains all the interdiction actions the leader can take to block some of the follower's structure choices.

For any action pair $(U, T) \in \Pi \times \Omega$, we will say that strategy U blocks structure T if $|U \cap T| \geq 1$, and define $\Omega_U = \{T \in \Omega \mid |U \cap T| \geq 1\}$ to be the collection of structures that are blocked by such a strategy. Consequently, if the leader chooses as their interdiction strategy the set U, the follower is then blocked from selecting any structure from Ω_U. Given a cost function $c : \Delta \to \mathbb{R}^+$ and a predefined threshold value $r \in \mathbb{R}$, the objective of the leader is to select a strategy $U \in \Pi$ of minimum cost such that the optimal choice of the follower is a structure whose weight is no less than r. In other words, the leader aims to minimize the cost of adopting an interdiction strategy that ensures a desired disruption level is inflicted on the follower's objective. This type of binary interdiction game can be modeled as follows

$$\begin{aligned} \min_{U \in \Pi} \quad & c(U) \\ \text{s.t.} \quad & \min_{T \in \Omega \setminus \Omega_U} w(T) \geq r. \end{aligned} \tag{1}$$

It is important to note that many interdiction games in the literature are modeled in terms of a limited budget: the leader's objective is set to maximize the weight of the optimal structure selected by the follower while ensuring that the cost of the chosen interdiction strategy is kept below a predefined budget b. In what follows, we will focus on binary interdiction games that can be formulated as (1) and then discuss an extension to address the budget version in Section 6.

Interestingly, considering the abstract nature of the structure and strategy sets in this model, it is no surprise that binary interdiction encapsulates a wide variety of adversarial games extensively studied in recent literature. For example, when the ground set Δ is composed of the edges or vertices of a given network, the aforementioned characterization can be used to model problems where the leader aims to restrict the follower's ability to conduct some operation in the network represented by structures such as shortest paths [32], spanning trees [60], cliques [23, 24, 42, 43], connected components [2, 18], matchings [67], dominating sets [46], or vertex covers [8]. Indeed, applications of such types of problems have found their niche in a wide variety of areas, including homeland security [26, 30, 31, 45, 63], disaster management [44], immunization strategies [56], sex trafficking prevention [64], energy systems [51], supply chain management [45], communications [19, 61], and transportation and logistics [33], among others.

Furthermore, recent interest in interdiction has spurred new developments, including several variations and extensions for these problems. Some notable examples include three-player interdiction games in which the new player, often called the protector, aims to conduct some preemptive actions to mitigate the effects of the interdiction strategy on the follower's structures [3, 13, 39]; two-player simultaneous interdiction games where both players act simultaneously without knowing in advance the strategy chosen by the other [28, 58]; stochastic interdiction games where some data of the problem follows some random probability distribution [16]; dynamic interdiction where the interaction between the two players is repeated through multiple rounds [52]; interdiction games with incomplete or asymmetric information where the knowledge of the underlying data is incomplete for one of the players or perceived differently [12, 49, 66]. Interested readers can refer to [53] for a general discussion about interdiction games, including solution methods and other variations.

Among the different solution methods, most exact approaches can be classified as either *dualize-and-combine* or *sampling/enumerative* methods [53]. The first type is typically reserved for cases where the structure set Ω admits a characterization in the form of a convex set (often, a polyhedron). In such cases, the inner optimization problem of the bi-level model in (1) can be replaced by its dual

representation. Then, utilizing strong duality and linearization techniques, it can be reformulated as a single-level optimization problem that can be solved by most off-the-shelf optimization solvers [16, 32, 37, 60]. Dualize-and-combine approaches like these mainly stem from developments in the broader class of bi-level optimization of which binary interdiction is a part of [5, 17, 35].

As for the sampling/enumerative algorithms, the main idea is to start with a small *sample* of structures in Ω that the leader would like to block and systematically identify new ones as the algorithm progresses. To this end, most approaches utilize mathematical models that are iteratively populated with new constraints that force the interdiction strategy to block the new structures that are progressively generated [15, 23, 32, 39, 43, 46, 60, 61].

For the particular case of binary interdiction games characterized by (1), the authors in [59] denote $\hat{\Omega} = \{T \in \Omega \mid w(T) < r\}$ as the set of "critical" structures that the leader should block; then, by letting $\mathbf{x} \in \{0,1\}^{|\Delta|}$ be the indicator vector of the leader's interdiction strategy and X_Π the characterization of the strategy set Π over the x space, they use the following model as a valid reformulation of (1)

$$\min \quad \sum_{a \in \Delta} c_a x_a \tag{2a}$$

$$\text{s.t.} \quad \sum_{a \in T} x_a \geq 1 \quad \forall T \in \hat{\Omega} \tag{2b}$$

$$\mathbf{x} \in X_\Pi. \tag{2c}$$

Here, for the sake of exposition, a linear cost function c is used, but note that the formulation can be directly adapted to accommodate other types of cost functions. Furthermore, in [59], the authors also proved that it is sufficient to use only the constraints in (2b) that are associated with the *minimal* critical structures in $\hat{\Omega}$, i.e., the critical structures that do not contain any strict subsets in $\hat{\Omega}$. Hence, we refer to this formulation as the minimal critical structures (MCS) formulation. We will subsequently refer to (2) by such a name too.

In recent years, several papers have introduced different families of inequalities to strengthen the mathematical formulations used to solve these problems [21, 23, 32, 37, 43, 46, 55, 60]. While some efforts have been focused on general cuts that can be applied to multiple types of interdiction games [15, 39, 60], most of the literature focuses on ad hoc inequalities tailored for some particular kind of structures [23, 46, 59]. In this paper, we are interested in a specific kind of constraints generally called *supervalid inequalities* [7, 32, 34, 50, 60]. In essence, a supervalid inequality is a constraint that can be added to strengthen a mathematical program such as (2) that may cut off feasible or even optimal solutions as long as a given incumbent solution already dominates them. Their usage in the context of interdiction stems from the seminal work by Israeli and Wood [32], where those are specifically designed to solve shortest-path interdiction games.

In the current literature, despite some promising computational results, these types of inequalities have only been developed sparingly for interdiction games with specific follower's structures, such as paths [32], vehicle routing plans [34], or spanning trees [60]. To the best of our knowledge, this paper is the first to study a general class of supervalid inequalities that can be applied to the broad class of binary interdiction games. Moreover, the nature of the supervalid inequalities studied in this paper is quite different from the current developments in the literature. We will compare our results with those in detail in Section 6. The contributions of this paper follow.

1. We identify a relationship between any bipartition $\mathcal{P}$ of the leader's strategy space Π (i.e., $\mathcal{P}$ is a partition of Π with two parts) and a collection of subsets from ground set Δ, denoted $\mathcal{P}$*-structures*, each of which induces a supervalid inequality for the MCS formulation. We

provide a general characterization of these inequalities and develop an algorithmic approach to use them in practice.

2. We study the separation process of the proposed supervalid inequalities. First, we develop an exact verification method used to identify whether any set $S \subseteq \Delta$ is a $\mathcal{P}$-structure. We show, however, that using such a verification method can be computationally challenging in practice. To improve the overall efficacy of this separation method, we then identify two types set constructions, based on which a more efficient verification becomes achievable. As a side product, all these characterizations also lead to a hierarchy map of the $\mathcal{P}$-structures.

3. We provide three general examples where we apply our results to derive supervalid inequalities for binary interdiction games targeting shortest paths, spanning trees, and vertex covers. In doing so, we also identify some interesting properties of the particular $\mathcal{P}$-structures associated with each of these problems.

4. Finally, we derive a connection between the $\mathcal{P}$-structures and a special type of set system called *greedoid* [36]. We prove that, under mild assumptions, the $\mathcal{P}$-structure separation procedure is guaranteed to be efficient for greedoid interdiction games.

The rest of the paper is organized as follows. In Section 2, we introduce some notation and three binary interdiction examples used throughout the paper for illustration purposes. In Section 3, we derive a correspondence between classes of supervalid inequalities and bipartitions of the leader's strategy space. In Section 4, we develop several methods to separate the proposed supervalid inequalities. In Section 5, we explore the connection between the proposed supervalid inequalities and a particular type of set system called a greedoid. In Section 6, we briefly discuss the application of the proposed method in interdiction games with a limited budget. In Section 7, we discuss some concluding remarks. Finally, we provide a computational demonstration of the proposed developments as part of the Appendix.

2 Preliminaries

2.1 Notation

We now proceed to introduce some notation that will be used throughout the paper. A set system $(\Delta, \mathcal{K})$ is defined as a pair composed of a ground set Δ and a family of subsets $\mathcal{K} \subseteq 2^{\Delta}$, such that $\mathcal{K}$ is a *partially ordered set* (*poset*) under the inclusion ($\subseteq$). Some poset-related notions that will be used hereafter are listed below.

Definition 1. Given a poset $(\mathcal{K}, \subseteq)$,

- let $m(\mathcal{K})$ be the set of all minimal elements in $\mathcal{K}$.
- $\mathcal{S}$ is a subposet of $\mathcal{K}$ if $\mathcal{S}$ is a subset of $\mathcal{K}$ with the inherited ordering.
- $\mathcal{S}$ is a lower (or upper) set in $\mathcal{K}$ if $\mathcal{S}$ is a subposet of $\mathcal{K}$ and, for any $U_1, U_2 \in \mathcal{K}$ such that $U_1 \subseteq U_2$ (or $U_1 \supseteq U_2$), $U_2 \in \mathcal{S}$ implies that $U_1 \in \mathcal{S}$.
- Given a subset $\mathcal{U}$ of a poset $\mathcal{K}$,

  $$\uparrow \mathcal{U} = \{U' \in \mathcal{K} \mid U' \supseteq U \text{ for some } U \in \mathcal{U}\} \text{ and } \downarrow \mathcal{U} = \{U' \in \mathcal{K} \mid U' \subseteq U \text{ for some } U \in \mathcal{U}\}$$

  are called the upper and lower closure of $\mathcal{U}$ in $\mathcal{K}$, respectively. Throughout the paper, we use $\uparrow \mathcal{U}$ ($\downarrow \mathcal{U}$) to denote the upper (lower) closure with respect to $\mathcal{K} = 2^{\Delta}$.

2.2 Examples of Binary Interdiction Games

In this subsection, we introduce three well-known interdiction games that fit into the binary characterization described before. They will be used for illustrative purposes throughout the paper.

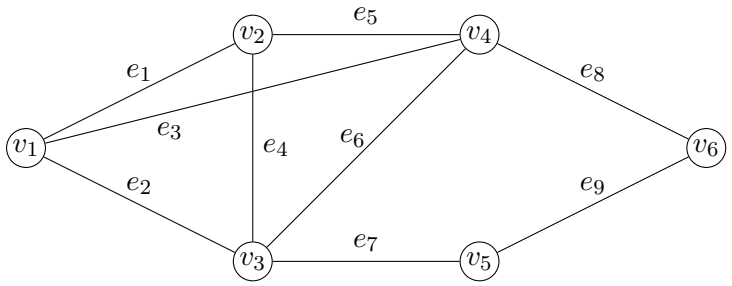


Figure 1: Example of a graph $G = (V, E)$ with $V = \{v_1, \ldots, v_6\}$ and $E = \{e_1, \ldots, e_9\}$.

Example 1 (Spanning Tree Interdiction [60])**.** Given an undirected graph $G = (V, E)$, with edge weights $\{w_e\}_{e \in E}$ and edge costs $\{c_e\}_{e \in E}$, a *spanning tree* is a connected acyclic subgraph that spans the vertex set V (e.g., $\{e_1, e_2, e_5, e_7, e_8\}$ in Figure 1). Let r be an interdiction target value; then, the spanning tree interdiction problem seeks to identify a subset of edges $U \subseteq E$ of minimum cost $\sum_{e \in U} c_e$, so that when removed from G, the minimum weight of any spanning tree left in the graph is at least r. Here, if the interdiction strategy U disconnects the graph, no spanning tree can be selected by the follower. Thus, by convention, its objective is assumed to be infinity. In this problem, Π contains all the edge sets and Ω is the set of spanning trees. $\triangle$

Example 2 (Shortest Path Interdiction [32])**.** Given a directed graph $G = (V, E)$, with edge costs $\{c_e\}_{e \in E}$ and edge lengths $\{w_e\}_{e \in E}$, two terminal vertices $s, t \in V$, and an interdiction target r, the shortest path interdiction problem aims to identify a set of edges $U \subseteq E$ of minimum cost $\sum_{e \in U} c_e$ so that when removed from G, the length of the shortest s-t paths that are left is at least r. As before, if the interdiction strategy U disconnects terminals s and t, the follower's objective is assumed to be infinity. Here, Π contains all the edge sets and Ω is the set of paths. $\triangle$

Example 3 (Vertex Cover Interdiction [9])**.** Given a graph $G = (V, E)$, with vertex costs $\{c_i\}_{i \in V}$ and vertex weights $\{w_i\}_{i \in V}$, a vertex cover is a vertex subset $D \subseteq V$ such that all edges in G are incident to at least one vertex in D (e.g., $D = \{v_1, v_2, v_4, v_5\}$ in Figure 1). Given an interdiction target r, the vertex cover interdiction problem is to identify a set of vertices $U \subseteq V$ of minimum costs to be blocked so that the weight of any vertex cover $D \subseteq V \setminus U$ of graph G is at least r. Similarly, if no vertex cover can be formed by selecting vertices from $D \subseteq V \setminus U$, the follower's objective is assumed to be infinity. In this problem, Π contains all the vertex sets and Ω is the set of vertex covers. $\triangle$

3 Bipartition-Induced Supervalid Inequalities

3.1 Motivation

Many interdiction problems share an interesting feature: when the search for a solution is conducted over some particular restriction of the leader's solution space, the resulting problem often becomes significantly easier to solve. To illustrate this, consider the shortest path interdiction game depicted in Figure 2. Here, among the five possible s-t paths available for the follower, we will assume that

the first three are critical—i.e., the lengths of such paths are less than the interdiction target r. Under these settings, the leader's strategy set $\Pi = 2^E$ contains all the possible edge subsets in the graph, and among the feasible solutions (i.e., edge sets in Π that intersect all critical structures in $\hat{\Omega}$), the minimal strategies that block all the critical paths are the following four edge sets

$$\{e_1, e_2, e_5\}, \{e_3, e_4, e_5\}, \{e_1, e_3, e_5\}, \{e_2, e_4, e_5\}.$$

Notice that the first two strategies are somewhat different from the other two as they are minimal s-t cuts. In fact, since any s-t cut is a feasible solution for the leader, one can use this *property* to produce the following partition of the leader's solution space Π:

$$\Pi_1 := \uparrow \{\{e_1, e_2, e_5\}, \{e_3, e_4, e_5\}\}, \quad \Pi_0 := \Pi \setminus \Pi_1.$$

Here, observe that set Π_1 is composed of the two strategies mentioned before and their supersets, as those are all the s-t cuts of the given graph. Furthermore, considering that finding an s-t cut of minimum cost can be done in polynomial time [29], solving the shortest path interdiction problem over the restricted space Π_1 is rather easy since it is equivalent to finding a minimum s-t cut, whereas in general solving the problem over Π is NP-hard [4].

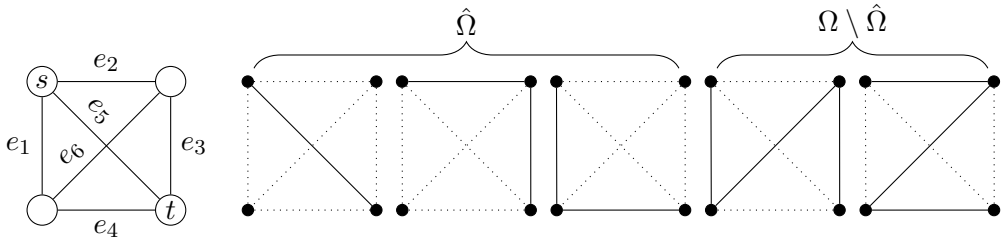


Figure 2: Illustration of bipartition in the shortest path interdiction.

This phenomenon seems to be a common occurrence in binary interdiction. Indeed, for the minimum spanning tree case, it is easy to see that any (global) edge-cut is a feasible strategy to block all spanning trees of a graph. Moreover, since finding a minimum cost edge-cut can be done in polynomial time [54], restricting the search over the set Π_1 composed of all edge cuts in the graph is a simple task while directly searching for an optimal solution over Π is, in general, NP-hard [10, 22].

In light of this discussion, a reasonable strategy when solving binary interdiction games is to first identify the best solution out of Π_1 and then focus all efforts on the more "difficult" task of searching for a potentially better solution over the remaining set Π_0. To speed up this search, one may wonder if it is possible to use relevant information about the given bipartition $\{\Pi_0, \Pi_1\}$ to produce a tighter mathematical formulation focused on searching for solutions in Π_0 rather than over the entire solution space Π. The goal is to systematically derive inequalities for (2) that are valid for the feasible solutions in Π_0 but may cut off strategies from Π_1, as well as some undesired fractional solutions from the formulation's original LP relaxation. The expected result is a tighter formulation defined by *supervalid* inequalities.

The use of exploratory searches over restricted solution spaces to speed up the running times of exact optimization solvers has been a common practice in integer programming since its conception. For example, exact approaches based on branch and bound often use local-search-based algorithms to explore some predefined neighborhood to identify good candidate solutions that may help prune unpromising branches [62]. While this type of approach can be directly applied here as well, our

aim goes beyond using the search over Π_1 to heuristically generate candidate solutions. Instead, we focus on producing stronger mathematical formulations.

As will be discussed in the following sections, our approach relies on exploiting a bipartition of the solution space Π into a set Π_1 whose exploration is "easy" and a set Π_0 whose exploration is "difficult". An important remark is that here we are using the terms *easy/difficult* rather loosely, as our methodology works for any bipartition of the solution space Π—even for cases where finding the best solution in Π_1 is also NP-Hard. However, in practice, our method is better suited for bipartitions for which solving the problem over Π_1 can be done in polynomial time. Furthermore, considering that the characterization of set Π_1 is generally linked to some structural attributes of its members (e.g., s-t cuts in a graph), in what follows, we will assume that the bipartitions $\{\Pi_0, \Pi_1\}$ are induced by some given property.

3.2 Bipartition Property $\mathcal{P}$ and the $\mathcal{P}$-Structure Formulation

We begin with the following definitions.

Definition 2 (Bipartition Property)**.** Given a set Π, a *bipartition property* is a function $\mathcal{P} : \Pi \to \{0, 1\}$. We set $\Pi_i^{\mathcal{P}} = \{U \in \Pi \mid \mathcal{P}(U) = i\}$ for $i \in \{0, 1\}$. When $\mathcal{P}$ is clear from the context, we drop the superscript. Given two bipartition properties $\mathcal{P}_1$ and $\mathcal{P}_2$, we say $\mathcal{P}_2$ is stronger than $\mathcal{P}_1$, denoted $\mathcal{P}_1 \preceq \mathcal{P}_2$, if and only if $\Pi_1^{\mathcal{P}_1} \subseteq \Pi_1^{\mathcal{P}_2}$.

Definition 3 (Cut/Cut Operator)**.** Given a family of sets $\Pi \subseteq 2^{\Delta}$, $S \subseteq \Delta$ is a *cut* of Π if $|S \cap U| \geq 1$ for all $U \in \Pi$. The *cut operator* $\mathcal{C} : 2^{\Delta} \to 2^{\Delta}$ is defined as $\mathcal{C}(\Pi) = \{S \subseteq \Delta \mid S \text{ is a cut of } \Pi\}$.

Notice that the cut operator also establishes a connection between the leader's feasible strategies, denoted by $\hat{\Pi}$, and the follower's critical structures $\hat{\Omega}$ through the following identity, which can be verified directly from the definitions

$$\hat{\Pi} = \mathcal{C}(\hat{\Omega}) \cap \Pi.$$

In fact, it is easy to see that each constraint (2b) of the MCS formulation is there to enforce that the feasible solutions of the leader represent cuts of $\hat{\Omega}$.

Some properties of the cut operator $\mathcal{C}$ are listed in the following proposition,

Proposition 1. *Given two posets $\mathcal{K}, \mathcal{K}' \subseteq 2^{\Delta}$ with the ordering inherited from $\subseteq$ on 2^{Δ}, the following statements about the cut operator $\mathcal{C}$ are true.*

1. *$\mathcal{C}(\emptyset) = 2^{\Delta}$ and $\mathcal{C}(2^{\Delta}) = \emptyset$.*
2. *$\mathcal{C}(\mathcal{K}) = \emptyset$ if $\emptyset \in \mathcal{K}$.*
3. *$\mathcal{C}(\mathcal{K})$ is an upper set.*
4. *$\mathcal{C}(\mathcal{K}) = \mathcal{C}(m(\mathcal{K}))$.*
5. *$\mathcal{C}$ is decreasing, i.e., if $\mathcal{K} \subseteq \mathcal{K}'$, then $\mathcal{C}(\mathcal{K}) \supseteq \mathcal{C}(\mathcal{K}')$.*
6. *$\mathcal{C}(\mathcal{C}(\mathcal{K})) = \uparrow \mathcal{K}$.*

Proof. For 1, the first part is vacuously true, and the second part is true because $\emptyset \in 2^{\Delta}$. Statements 2-5 can be verified directly from the definition of $\mathcal{C}$. For 6, if $\emptyset \in \mathcal{K}$, then by Statements 1 and 2, $\mathcal{C}(\mathcal{C}(\mathcal{K})) = 2^{\Delta}$, which is equal to $\uparrow \mathcal{K}$ in 2^{Δ}, since $\emptyset \in \mathcal{K}$. Suppose $\emptyset \notin \mathcal{K}$, then we have,

$$U \in \mathcal{C}(\mathcal{C}(\mathcal{K})) \iff \forall S \in \mathcal{C}(\mathcal{K}),\ |U \cap S| \geq 1 \iff U \in \uparrow \mathcal{K},$$

where the first equivalence follows from the definition of the operator $\mathcal{C}$. For the "$\Leftarrow$" direction of the second equivalence, $U \in\uparrow \mathcal{K}$ means $U \supseteq U'$ for some $U' \in \mathcal{K}$. By definition, for all $S \in \mathcal{C}(\mathcal{K})$, $|S \cap U'| \geq 1$, then this is also true for the superset U. For "$\Rightarrow$", we prove the contrapositive. Given $U \notin\uparrow \mathcal{K}$ and $\emptyset \notin \mathcal{K}$ by assumption, then for every $U' \in \mathcal{K}$, $U' \setminus U \neq \emptyset$. Then, $S = \bigcup_{U' \in \mathcal{K}} U' \setminus U$ intersects all elements in $\mathcal{K}$, but $S \cap U = \emptyset$ by construction. This negates the second sentence in the above equivalence chain, completing the proof. □

Using the cut operator, we define the following structures that are induced by a given bipartition property $\mathcal{P}$.

Definition 4 ($\mathcal{P}$-Structure)**.** Given the set of follower's critical structures $\hat{\Omega} \subseteq 2^{\Delta}$, the set of leader's strategies Π, and a bipartition property $\mathcal{P} : \Pi \to \{0, 1\}$, the set of bipartition-induced structures with respect to $\mathcal{P}$, or simply $\mathcal{P}$-structures, is defined as

$$\hat{\Omega}_{\mathcal{P}} := \mathcal{C}(\mathcal{C}(\hat{\Omega}) \cap \Pi_0) = \mathcal{C}(\hat{\Pi} \cap \Pi_0).$$

Since $\hat{\Pi} = \mathcal{C}(\hat{\Omega})$ is the set of all the feasible strategies, we use $\hat{\Pi}_0$ to denote all the feasible strategies in Π_0. Hence, $\hat{\Omega}_{\mathcal{P}} = \mathcal{C}(\hat{\Pi}_0)$.

Intuitively, a $\mathcal{P}$-structure is a subset of the ground set Δ that intersects all the feasible strategies in $\hat{\Pi}_0$ that resulted from the bipartition property $\mathcal{P}$. These $\mathcal{P}$-structures are particularly important for our development, as we will soon show that they can directly induce supervalid inequalities for Formulation (2).

Since every bipartition property $\mathcal{P}$ induces a certain set of $\mathcal{P}$-structures, it is then reasonable to analyze first the following two extreme properties.

Proposition 2. *Define bipartition properties* $\mathbf{0}$ *and* $\mathbf{1}$ *as*

$$\mathbf{0}(U) = 0 \text{ and } \mathbf{1}(U) = 1$$

for all $U \in \Pi$*. Then, we have* $\hat{\Omega}_{\mathbf{0}} =\uparrow \hat{\Omega}$ *and* $\hat{\Omega}_{\mathbf{1}} =\uparrow \emptyset = 2^{\Delta}$.

Proof. For $\mathbf{0}$, we have $\Pi_0 = \Pi$, which leads to the following

$$\hat{\Omega}_{\mathbf{0}} = \mathcal{C}(\hat{\Pi}) = \mathcal{C}(\mathcal{C}(\hat{\Omega})) =\uparrow \hat{\Omega},$$

where the first two identities are by Definition 4 and the last is by the statement *6.* of Proposition 1. A similar verification can be done for property $\mathbf{1}$. □

Intuitively, these two extreme properties correspond to the following two cases. First, the $\mathbf{0}$ property induces the bipartition where set $\Pi_1 = \emptyset$, thus all leader strategies are contained in $\Pi_0 = \Pi$. Here, since the problem over Π_0 is the same as the problem over Π, the corresponding $\mathcal{P}$-structure set is simply the upper closure of the critical structures in $\hat{\Omega}$. Second, the $\mathbf{1}$ property induces the bipartition where $\Pi_1 = \Pi$ and thus $\Pi_0 = \emptyset$. Here, under the assumption that Π_1 is "easy to explore", property $\mathbf{1}$ represents the extreme case where the original interdiction problem is solvable in polynomial time. Any property $\mathcal{P}$ that is strictly sandwiched between $\mathbf{0}$ and $\mathbf{1}$ (i.e., $\mathbf{0} \preceq \mathcal{P} \preceq \mathbf{1}$) will produce nontrivial $\mathcal{P}$-structures. We are interested in these $\mathcal{P}$-structures due to the following theorem.

Theorem 1. *In a binary interdiction game with leader's strategy set* Π *and follower's critical structure set* $\hat{\Omega}$*, given any bipartition property* $\mathcal{P}$ *on* Π*, an optimal strategy* $U^\star$ *must satisfy one of the following two assertions,*

- $U^\star \in \Pi_1$;
- $U^\star \in \mathcal{C}(\hat{\Omega}_\mathcal{P})$.

Proof. An optimal strategy $U^\star$ that is not in Π_1 must be in $\hat{\Pi}_0$ since Π_0 and Π_1 form a bipartition of Π and $U^\star$ is feasible. On the other hand, we have

$$\mathcal{C}(\hat{\Omega}_\mathcal{P}) = \mathcal{C}(\mathcal{C}(\hat{\Pi}_0)) =\uparrow \hat{\Pi}_0 \supseteq \hat{\Pi}_0,$$

which concludes the proof. □

This theorem allows us to utilize a bipartition on Π and the assumed "easy-to-obtain" solution in Π_1 to set up a stronger formulation for the problem defined over the "difficult-to-explore" space Π_0. In particular, we get the following minimal $\mathcal{P}$-structure (MPS) formulation,

$$\begin{aligned}
\text{(MPS)} \quad \min \quad & \sum_{a\in\Delta} c_a x_a && \text{(3a)}\\
\text{s.t.} \quad & \sum_{a\in T} x_a \geq 1 \quad \forall T \in m(\hat{\Omega}_\mathcal{P}) && \text{(3b)}\\
& \mathbf{x} \in X_\Pi. && \text{(3c)}
\end{aligned}$$

Note that this formulation only searches for an optimal strategy within $\uparrow \Pi_0$, thus it has a smaller solution space than the MCS formulation. In other words, the inequalities in (3b) are indeed supervalid for the general space Π, as they may remove feasible solutions to the original problem. The following corollary characterizes the solution space of (3) exactly.

Corollary 1. *A strategy $U \in \Pi$ is feasible to the MPS formulation if and only if it is a superset of some $U' \in \hat{\Pi}_0$.*

Proof. For sufficiency, note that any $U' \in \hat{\Pi}_0$ intersects all elements in $\hat{\Omega}_\mathcal{P}$ by definition, so as any superset of U'. For necessity, we assume $\emptyset \notin \hat{\Pi}_0$, otherwise, all strategies are supersets of $\emptyset$. Hence, to prove the contrapositive, given a strategy U that is not a superset of any $U' \in \hat{\Pi}_0$, we have $U' \setminus U \neq \emptyset$ for all $U' \in \hat{\Pi}_0$. Let $S = \bigcup_{U'\in\hat{\Pi}_0} U' \setminus U$, then S is a cut of $\hat{\Pi}_0$ but $S \cap U = \emptyset$ by construction. The former implies $S \in \hat{\Omega}_\mathcal{P}$ and the latter means U does not intersect the $\mathcal{P}$-structure S, that is, U is not feasible to the MPS formulation. □

The following corollary further analyzes the effect of choosing different bipartition properties.

Corollary 2. *Given two bipartition properties $\mathcal{P}$ and $\mathcal{P}'$, if $\mathcal{P} \preceq \mathcal{P}'$ then $\hat{\Omega}_\mathcal{P} \subseteq \hat{\Omega}_{\mathcal{P}'}$. In particular, $\hat{\Omega} \subseteq \hat{\Omega}_\mathbf{0} \subseteq \hat{\Omega}_\mathcal{P}$ for every bipartition property $\mathcal{P}$.*

Proof. By definition, $\mathcal{P} \preceq \mathcal{P}'$ if and only if $\Pi_0^\mathcal{P} \supseteq \Pi_0^{\mathcal{P}'}$, which implies $\hat{\Pi}_0^\mathcal{P} \supseteq \hat{\Pi}_0^{\mathcal{P}'}$. Then, the claim is true since the operator $\mathcal{C}$ is decreasing by the statement *5.* of Proposition 1. □

This result confirms the intuition that the constraint set induced by $\hat{\Omega}_\mathcal{P}$ becomes stronger when the $\mathcal{P}$ carves out a larger portion of the solution space Π into the "easy-to-explore" part Π_1.

Next, we will provide three concrete examples to illustrate the corresponding $\mathcal{P}$-structures.

3.3 Examples of $\mathcal{P}$-Structures

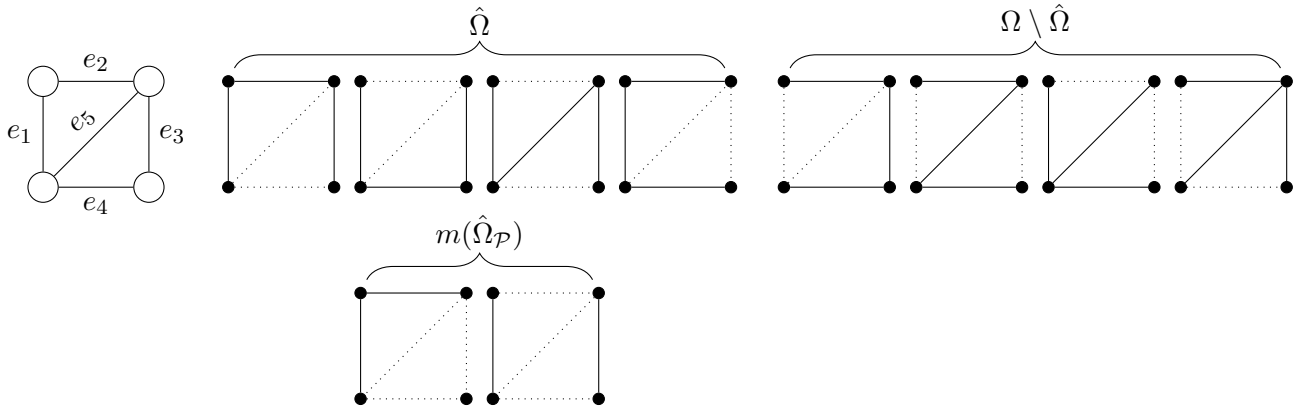
Example 1 (Spanning Tree Interdiction, Cont.)**.** As discussed previously, a natural bipartition property $\mathcal{P}$ for the minimum spanning tree interdiction is whether a solution is an edge-cut of graph G, since the edge cut of minimum cost can be found in polynomial time [54]. That is, for all $U \in \Pi$,

$$\mathcal{P}(U) = \begin{cases} 1, & U \text{ is an edge-cut,} \\ 0, & \text{otherwise.} \end{cases}$$

Figure 3 provides an example of the corresponding $\mathcal{P}$-structures for the minimum spanning tree interdiction problem over a 4-vertex graph. Here, among the eight spanning trees comprising the follower's structure set Ω, we assume the first four are the critical ones. The computation of $m(\hat{\Omega}_{\mathcal{P}})$ follows,

- $\hat{\Omega} = \{\{e_1, e_2, e_3\}, \{e_1, e_3, e_4\}, \{e_1, e_3, e_5\}, \{e_1, e_2, e_4\}\}$.
- $m(\hat{\Pi}) = \{\{e_1\}, \{e_2, e_3\}, \{e_3, e_4\}, \{e_2, e_4, e_5\}\}$.
- $m(\hat{\Pi}_0) = \{\{e_1\}, \{e_2, e_3\}\}$, since the other two are edge-cuts.
- $m(\hat{\Omega}_{\mathcal{P}}) = m(\mathcal{C}(m(\hat{\Pi}_0))) = \{\{e_1, e_2\}, \{e_1, e_3\}\}$.

This leads to the corresponding MPS formulation in Figure 3. Notice that the feasible edge cut solution $\{e_3, e_4\}$ becomes infeasible to this MPS formulation as it violates the first constraint. This confirms that these inequalities are supervalid for Π. We also note that, in this example, every $\mathcal{P}$-structure is a subset of some critical spanning tree.



(MCS)
$$\begin{aligned} \min \quad & \mathbf{cx} \\ \text{s.t.} \quad & x_1 + x_2 + x_3 \geq 1 \\ & x_1 + x_3 + x_4 \geq 1 \\ & x_1 + x_3 + x_5 \geq 1 \\ & x_1 + x_2 + x_4 \geq 1 \\ & \mathbf{x} \in \{0,1\}^E \end{aligned}$$

(MPS)
$$\begin{aligned} \min \quad & \mathbf{cx} \\ \text{s.t.} \quad & x_1 + x_2 \geq 1 \\ & x_1 + x_3 \geq 1 \\ & \mathbf{x} \in \{0,1\}^E \end{aligned}$$

Figure 3: $\mathcal{P}$-structures and supervalid inequalities in the minimum spanning tree interdiction.

In [60], these minimal $\mathcal{P}$-structures are called *critical edge sets* and are derived from a different perspective. $\triangle$

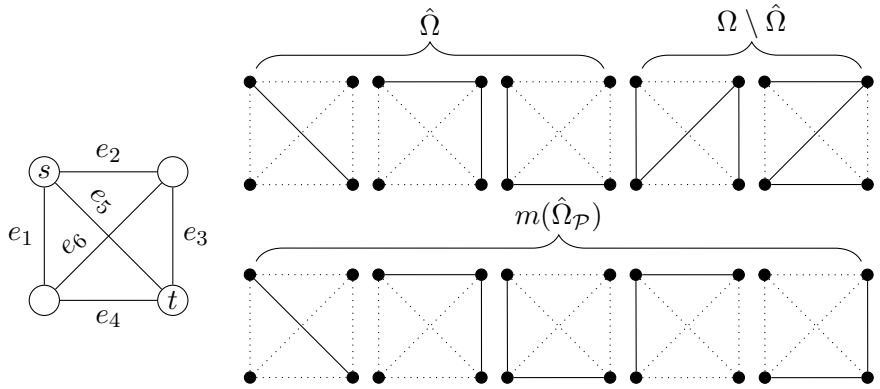
Example 2 (Shortest Path Interdiction, Cont.)**.** As mentioned before, a natural bipartition property for the shortest path interdiction is as follows. For each $U \in \Pi$,

$$\mathcal{P}(U) = \begin{cases} 1, & U \text{ is a } s\text{-}t \text{ cut}, \\ 0, & \text{otherwise}. \end{cases}$$

Figure 4 provides an example of the corresponding $\mathcal{P}$-structures for the shortest path interdiction problem over a 4-vertex graph. Here, among the five s-t paths comprising the follower's structure set Ω, we assume the first three are the critical ones. As before, we compute all the corresponding minimal $\mathcal{P}$-structures as follows.

- $\hat{\Omega} = \{\{e_1, e_4\}, \{e_2, e_3\}, \{e_5\}\}$.
- $m(\hat{\Pi}) = \{\{e_1, e_2, e_5\}, \{e_1, e_3, e_5\}, \{e_2, e_4, e_5\}, \{e_3, e_4, e_5\}\}$.
- $m(\hat{\Pi}_0) = \{\{e_1, e_3, e_5\}, \{e_2, e_4, e_5\}\}$, since the other two are s-t cuts.
- $m(\hat{\Omega}_{\mathcal{P}}) = m(\mathcal{C}(m(\hat{\Pi}_0))) = \{\{e_1, e_2\}, \{e_1, e_4\}, \{e_2, e_3\}, \{e_3, e_4\}, \{e_5\}\}$.

This generates the corresponding MPS formulation in Figure 4. Again, the feasible s-t cut solutions $\{e_1, e_2, e_5\}$ and $\{e_3, e_4, e_5\}$ violate constraints $x_3 + x_4 \geq 1$ and $x_1 + x_2 \geq 1$, respectively, in the MPS formulation.



(MCS)
$$\begin{aligned} \min \quad & \mathbf{cx} \\ \text{s.t.} \quad & x_5 \geq 1 \\ & x_2 + x_3 \geq 1 \\ & x_1 + x_4 \geq 1 \\ & \mathbf{x} \in \{0,1\}^E \end{aligned}$$

(MPS)
$$\begin{aligned} \min \quad & \mathbf{cx} \\ \text{s.t.} \quad & x_5 \geq 1 \\ & x_2 + x_3 \geq 1 \\ & x_1 + x_4 \geq 1 \\ & x_1 + x_2 \geq 1 \\ & x_3 + x_4 \geq 1 \\ & \mathbf{x} \in \{0,1\}^E \end{aligned}$$

Figure 4: $\mathcal{P}$-structures and supervalid inequalities in the shortest path interdiction.

In this example, all the minimal critical s-t paths are also minimal $\mathcal{P}$-structures. One particular difference with respect to the minimum spanning tree case is that here set $m(\hat{\Omega}_{\mathcal{P}})$contains two extra structures—$\{e_1, e_2\}$ and $\{e_3, e_4\}$—that are not subsets of any s-t path. △

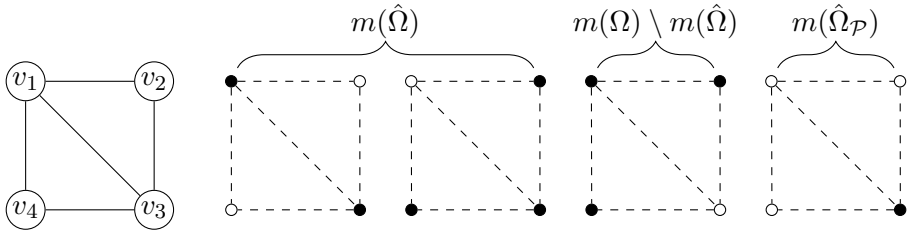
Example 3 (Vertex Cover Interdiction, Cont.)**.** By definition, each edge in a graph $G = (V, E)$ has at least one endpoint in any vertex cover $T \subseteq V$. Thus, any vertex set $U \in \Pi$ that contains two adjacent vertices blocks all the vertex covers in G. Therefore, any feasible solution that does not satisfy this property must be an independent set. Following this logic, a bipartition property for the vertex cover interdiction can be defined as follows. For each $U \in \Pi$,

$$\mathcal{P}(U) = \begin{cases} 0, & U \text{ is an independent set,} \\ 1, & \text{otherwise.} \end{cases}$$

Figure 5 provides an example of the corresponding $\mathcal{P}$-structures for the vertex cover interdiction problem over a 4-vertex graph. Here, it is easy to verify that there are six vertex covers $\{\{v_1, v_3\}, \{v_2, v_3, v_4\}, \{v_1, v_2, v_4\}, \{v_1, v_2, v_3\}, \{v_1, v_3, v_4\}, \{v_1, v_2, v_3, v_4\}\}$ comprising the follower's set Ω. We assume the first two are the critical ones. For convenience, in the figure, we only list the minimal vertex covers (in black). As before, we compute all the corresponding minimal $\mathcal{P}$-structures as follows.

- $\hat{\Omega} = \{\{v_1, v_3\}, \{v_2, v_3, v_4\}\}$.
- $m(\hat{\Pi}) = \{\{v_3\}, \{v_1, v_2\}, \{v_1, v_4\}\}$.
- $m(\hat{\Pi}_0) = \{\{v_3\}\}$, since the other two are not independent sets.
- $m(\hat{\Omega}_{\mathcal{P}}) = m(\mathcal{C}(m(\hat{\Pi}_0))) = \{\{v_3\}\}$.

This generates the corresponding MPS formulation in Figure 5. Again, the two feasible solutions $\{v_1, v_2\}$ and $\{v_1, v_4\}$ are removed by the supervalid constraints in the MPS formulation. Similarly to the minimum spanning tree example, every $\mathcal{P}$-structure here is a subset of some critical vertex cover. △



(MCS)
$$\begin{aligned} \min \quad & \mathbf{cx} \\ \text{s.t.} \quad & x_1 + x_3 \geq 1 \\ & x_2 + x_3 + x_4 \geq 1 \\ & \mathbf{x} \in \{0,1\}^V \end{aligned}$$

(MPS)
$$\begin{aligned} \min \quad & \mathbf{cx} \\ \text{s.t.} \quad & x_3 \geq 1 \\ & \mathbf{x} \in \{0,1\}^V \end{aligned}$$

Figure 5: $\mathcal{P}$-structures and supervalid inequalities in vertex cover interdiction.

In the next subsection, we will introduce an algorithm to solve the MPS formulation and prove its correctness.

Algorithm 1: MPS Algorithm.

Data: input $\mathbf{c}, \mathcal{P}, \Pi, \hat{\Omega}$

1 $\bar{z}, \bar{x} \leftarrow$ `EasySolve(`Π_1`)` `// get the optimal value and solution on` Π_1

2 $\bar{\Omega}_{\mathcal{P}} \leftarrow \emptyset$ `// initialize with an empty set of` $\mathcal{P}$`-structures`

3 $\underline{z} \leftarrow -\infty,\ \underline{x} \leftarrow \mathbf{0}$ `// initialize the lower bound`

4 **while** *true* **do**

5 $\quad z, x \leftarrow$ `MPS(`$\mathbf{c}$, Π, $\bar{\Omega}_{\mathcal{P}}$`)` `// solve the MPS formulation with the current` $\bar{\Omega}_{\mathcal{P}}$

6 $\quad$ **if** $z \geq \bar{z}$ **then**

7 $\quad\quad z^\star \leftarrow \bar{z},\ x^\star \leftarrow \bar{x}$ `// stop due to worse than` Π_1

8 $\quad\quad$ **break**

9 $\quad$ **else**

10 $\quad\quad \bar{\Omega}_x \leftarrow$ `MPSeparation(`$\mathbf{c}, \Pi, x$`)`

11 $\quad\quad$ **if** $\bar{\Omega}_x = \emptyset$ **then**

12 $\quad\quad\quad z^\star \leftarrow z,\ x^\star \leftarrow x$ `// stop due to optimality on` Π_0

13 $\quad\quad\quad$ **break**

14 $\quad\quad$ **else**

15 $\quad\quad\quad \bar{\Omega}_{\mathcal{P}} \leftarrow \bar{\Omega}_{\mathcal{P}} \cup \bar{\Omega}_x$ `// update the set of` $\mathcal{P}$`-structures`

16 $\quad\quad\quad \underline{z} \leftarrow z,\ \underline{x} \leftarrow x$ `// update the lower bound`

17 $\quad\quad$ **end**

18 $\quad$ **end**

19 **end**

20 **return** $z^\star, x^\star$

3.4 MPS Algorithm

Since the MPS formulation only searches for an optimal solution from Π_0, we need to combine the results from Π_1 to obtain the global optimal solution for (2). We call this combined procedure the MPS algorithm (Algorithm 1). There are three subroutines invoked in this algorithm: (i) the EasySolve subroutine that solves the problem over the "easy-to-explore" solution space Π_1 and obtains the corresponding optimal solution therein; (ii) the MPS subroutine that solves a relaxed version of the MPS formulation defined over a set $\bar{\Omega}_{\mathcal{P}} \subseteq \Omega_{\mathcal{P}}$; and (iii) the MPSeparation subroutine that iteratively updates $\bar{\Omega}_{\mathcal{P}}$ by identifying new $\mathcal{P}$-structures in $\Omega_{\mathcal{P}} \setminus \bar{\Omega}_{\mathcal{P}}$. This iterative procedure terminates when either no new $\mathcal{P}$-structures can be produced or the objective value from MPS is at greater than or equal to the solution obtained by EasySolve. The following theorem proves the correctness and finite termination of Algorithm 1.

Theorem 2. *The MPS algorithm terminates in finite iterations, and the returned solution $x^\star$ is optimal with respect to the MCS formulation.*

Proof. We first show that the number of iterations is finite. Let z_i and $\bar{\Omega}^i_{\mathcal{P}}$ denote the value of z and contents of $\bar{\Omega}_{\mathcal{P}}$ at Step 5 of the algorithm's ith iteration. Because of Step 15, we clearly have $z_i \leq z_{i+1}$ and $\bar{\Omega}^i_{\mathcal{P}} \subseteq \bar{\Omega}^{i+1}_{\mathcal{P}}$. Moreover, the latter inclusion is strict since, otherwise, we have $\bar{\Omega}_x = \emptyset$, and the procedure terminates due to Step 11–13. Since the set of $\mathcal{P}$-structures is finite, we have two possible cases: (i) at a certain iteration we get $\bar{\Omega}_{\mathcal{P}} = \hat{\Omega}_{\mathcal{P}}$ or (ii) the algorithm halts before (i) happens. For the former case, we have $\bar{\Omega}_x = \emptyset$ at the corresponding iteration, and, again, the algorithm terminates by Step 11–13, proving the claim.

For correctness, we again focus on the two cases introduced before. For (i), we have generated the full MPS formulation and have $z^\star < \bar{z}$. By Corollary 1, $x^\star$ is an optimal solution in Π_0.

Moreover, $x^\star$ is better than the optimal solution in Π_1 as $z^\star < \bar{z}$. This proves $x^\star$ is an optimal solution over the entire solution space Π of the MCS formulation. For Case (ii), the algorithm may terminate in two different ways: $(ii.1)$ $z \geq \bar{z}$; $(ii.2)$ $\bar{\Omega}_x = \emptyset$. Case (ii.1) implies the optimal solution in Π_0 has an objective value that is greater than or equal to the optimal value in Π_1, which proves the optimality of $x^\star = \bar{x}$. Case $(ii.2)$ means that solution x obtained at Step 5 is feasible for the MPS formulation and is optimal for a relaxation of the MPS formulation. Thus, x is also optimal to the MPS formulation, i.e., it is an optimal solution in Π_0. Again, because of $z < \bar{z}$, the solution x is optimal over the entire solution space Π. □

By assumption, subroutine EasySolve can be computed efficiently, and by Theorem 1 and Corollary 2, the MPS formulation has a strictly smaller solution space than the MCS formulation for any nontrivial bipartition property $\mathcal{P}$. Thus, the overall efficiency of the MPS algorithm depends on the trade-off between the strength of the constraints induced by the $\mathcal{P}$-structures and the complexity of the MPSeparation subroutine. We will devote the next section to the general theory of $\mathcal{P}$-structure separation.

4 $\mathcal{P}$-Structure Constraints Separation

Compared to the process of separating inequalities associated with the critical structures in $m(\hat{\Omega})$ (i.e., constraints (2b) of the MCS formulation), separating the constraints (3b) associated with a given set of $\mathcal{P}$-structures can be significantly harder. The main reason is that the elements in $m(\hat{\Omega})$ are feasible solutions for the follower that can be generated via the follower's optimization problem, whereas the $\mathcal{P}$-structures are arbitrary subsets of Δ without a direct method to produce them.

In this section, we focus on developing several methods for verifying whether any given set $S \subseteq \Delta$ is a $\mathcal{P}$-structure, which will naturally lead to constraint separation methods. Specifically, given a set $S \subseteq \Delta$, a set of critical structures $\hat{\Omega}$, and a bipartition property $\mathcal{P}$, a *$\mathcal{P}$-structure verifier* is a subroutine that returns true if $S \in \hat{\Omega}_{\mathcal{P}}$ and false otherwise. We say the verifier is *partial* if it only verifies a subset of $\hat{\Omega}_{\mathcal{P}}$, i.e., it may return a false negative but not a false positive.

A naive verification can be done by following the definition directly, as it was done for the three examples given in Section 3.3. However, this verification process is impractical even for modest-sized problems, as it requires knowing the entire feasible solution space $\hat{\Pi}$. In what follows, we restrict our attention to a special class of bipartition properties defined as follows.

Definition 5 (Null/Regular Properties)**.** Given the leader's strategy set Π and the follower's structure set Ω, the null property $\mathcal{P}_\emptyset$ is defined as follows. For every $U \in \Pi$,

$$\mathcal{P}_\emptyset(U) = \begin{cases} 1, & U \in \mathcal{C}(\Omega), \\ 0, & \text{otherwise.} \end{cases}$$

We say a property $\mathcal{P}$ is *regular* if and only if $\mathcal{P} \succeq \mathcal{P}_\emptyset$.

Hence, a strategy U belongs to $\Pi_1^{\mathcal{P}_\emptyset}$ if and only if it nullifies (intersects) every follower's structure in Ω (not just the critical structures in $\hat{\Omega}$). In fact, all three examples presented in Section 3.3 use the null property as their bipartition properties. Indeed, in those examples, every edge-cut, s-t cut, and non-independent set intersects all the spanning trees, s-t paths, and vertex covers, respectively. We will soon show that the $\mathcal{P}$-structures associated with null properties have some convenient characteristics that can be exploited for constraint separation purposes. Before diving into the verification algorithms, we need to provide some notation first.

4.1 Extended Sets

Definition 6. Given a family of sets Ω, we define

- $\Omega[S,k] := \{T \in \Omega \mid |T \cap S| \geq |S| - k\}$.
- $\Omega[S] := \Omega[S,0] = \{T \in \Omega \mid T \supseteq S\}$.

We call $\Omega[S,k]$ the set Ω extended on S with degree k and call $\Omega[S]$ the set Ω extended on S (with degree 0).

We have the following properties for the extended sets.

Proposition 3. *For sets $S \subseteq S'$ and numbers $k \leq k' \in \mathbb{N}$, we have*

1. $\Omega[S,k] = \Omega$ *for every* $k \geq |S|$.
2. $\Omega[S,k] \subseteq \Omega[S,k']$.
3. $\Omega[S,k] \supseteq \Omega[S',k]$.

Proof. Statement 1 can be verified directly by definition. For 2, give a fixed set S, we have

$$T \in \Omega[S,k] \Longrightarrow |T \cap S| \geq |S| - k \geq |S| - k' \Longrightarrow T \in \Omega[S,k'].$$

For 3, we know $|S'| = |S' \setminus S| + |S|$ since S' is a superset of S, and $|T \cap S'| \leq |T \cap S| + |S' \setminus S|$ because $T \cap S'$ contains at most everything in $S' \setminus S$ besides $T \cap S$. Then, we have

$$\begin{aligned} & T \in \Omega[S',k] \\ \Longrightarrow\ & |T \cap S'| \geq |S'| - k \\ \Longrightarrow\ & |T \cap S| + |S' \setminus S| \geq |S' \setminus S| + |S| - k \\ \Longrightarrow\ & |T \cap S| \geq |S| - k \\ \Longrightarrow\ & T \in \Omega[S,k]. \end{aligned}$$

This completes the proof. □

4.2 An Exact Verifier for $\mathcal{P}_\emptyset$

When focusing on null property $\mathcal{P}_\emptyset$, the following definition characterizes the corresponding set of $\mathcal{P}_\emptyset$-structures from a different perspective.

Definition 7 (Critical Nucleus)**.** A set S is a *critical nucleus* if for every $T \in (\uparrow \Omega)[S]$, there exists some $T' \in \hat{\Omega}$ such that $T' \subseteq T$.

By definition, $(\uparrow \Omega)[S]$ is an extended subposet. So, a set S is a critical nucleus if every superset of S in $\uparrow \Omega$ (with respect to 2^Δ) also contains some critical structure. For instance, consider the edges set $S = \{e_1, e_2\}$ in the shortest path interdiction example shown in Figure 2. By definition, every edge set in $(\uparrow \Omega)[S]$ must contain S along with some $T \in \Omega$. Thus, it is sufficient to check whether every $T \in \Omega \setminus \hat{\Omega}$ together with S can contain some critical path. In particular, the critical paths $\{e_2, e_3\}$ and $\{e_1, e_4\}$ are contained in the respective $S \cup \{e_1, e_3, e_6\}$ and $S \cup \{e_2, e_4, e_5\}$, which renders S as a critical nucleus.

Theorem 3. *Given $\hat{\Omega} \neq \emptyset$, $S \in \hat{\Omega}_{\mathcal{P}_\emptyset}$ if and only if S is a critical nucleus.*

Proof. For sufficiency, towards a contradiction, suppose $S \notin \hat{\Omega}_{\mathcal{P}_\emptyset} = \mathcal{C}(\hat{\Pi}_0)$, then there exists some feasible strategy $U \in \hat{\Pi}_0$ such that $S \cap U = \emptyset$. By definition of $\mathcal{P}_\emptyset$, such U intersects all critical structures in $\hat{\Omega}$ but not all structures in Ω for otherwise, $U \in \Pi_1$. Hence, we can pick some $T \in \{T' \in \Omega \mid T' \cap U = \emptyset\}$, and we have $(T \cup S) \cap U = \emptyset$. Clearly, $T \cup S \in (\uparrow \Omega)[S]$, then the fact that S is a critical nucleus implies that we can find some critical structure $T' \in \hat{\Omega}$ such that $T' \subseteq T \cup S$. By construction, $T' \cap U = \emptyset$, which implies U is not a feasible strategy, a contradiction.

For necessity, suppose S is not a critical nucleus, i.e., there is some $T \in (\uparrow \Omega)[S]$ such that for all $T' \in \hat{\Omega}$, $T' \not\subseteq T$. Then, take $U = \bigcup_{T' \in \hat{\Omega}} T' \setminus T$, U intersects all the critical structures in $\hat{\Omega}$ hence is a feasible strategy. Moreover, T is from $(\uparrow \Omega)[S]$ implies T includes some structure $T_0 \in \Omega$, and $U \cap T_0 = \emptyset$ by construction. Thus, $\mathcal{P}_\emptyset(U) = 0$ as U does not block all the structures. So, we have $U \in \hat{\Pi}_0$ and $U \cap S = \emptyset$, since $S \subseteq T$ by our choice. This implies S is not a cut of $\hat{\Pi}_0$, thus is not a $\mathcal{P}_\emptyset$-structure. □

A direct implication of this theorem is the following corollary.

Corollary 3. *For every regular $\mathcal{P}$, $\hat{\Omega}_{\mathcal{P}}$ contains all the critical nuclei.*

Proof. Every regular $\mathcal{P}$ is stronger than $\mathcal{P}_\emptyset$. Thus, by Corollary 2, we have $\hat{\Omega}_{\mathcal{P}_\emptyset} \subseteq \hat{\Omega}_{\mathcal{P}}$. □

For every regular bipartition property $\mathcal{P}$, the above theorem and corollary provide a different (partial) characterization of the $\mathcal{P}$-structures that does not depend on the feasible solutions space $\hat{\Pi}$ of the MCS formulation. This enables us to develop verifiers based on this new characterization. In particular, we have the following exact verifier for $\mathcal{P}_\emptyset$-structures.

Theorem 4. *A set S is a critical nucleus if and only if the following formulation*

$$\max_{T \in m(\Omega)} \min_{\substack{T' \subseteq T \cup S \\ T' \in m(\Omega)}} w(T') \tag{4}$$

has an optimal value that is less than r.

Proof. It is sufficient to show that S is a critical nucleus if and only if for every $T \in m(\Omega)$, $T \cup S$ contains some critical structure $T' \in \hat{\Omega}$. The necessity is obvious as the set $T \cup S$ for some $T \in m(\Omega)$ belongs to $(\uparrow \Omega)[S]$. For sufficiency, we prove the contrapositive. Suppose S is not a critical nucleus, i.e., there exists some $T \in (\uparrow \Omega)[S]$ such that every defender's structure T' contained in $T \cup S$ is non-critical. Notice such T' always exists as we pick T from $(\uparrow \Omega)$. Moreover, either T' is minimal in Ω, or some subset of T' is. Without loss of generality, assume $T' \in m(\Omega)$, then by choice, $T' \cup S$ does not contain any critical structure, which completes the proof. □

According to this theorem, Formulation (4) provides an exact verifier for the $\mathcal{P}_\emptyset$-structures. Since (4) is a bilevel problem, optimization techniques such as dualize-and-merge and branch-and-bound can be applied to verify a given S exactly. We provide below several examples to illustrate this method.

Example 1 (Spanning Tree Interdiction, Cont.)**.** In this problem, $m(\Omega)$ is simply the set of spanning trees. Thus, we can use the established spanning tree formulations [41] to represent the constraints of Formulation (4). For the outer level problem, we can use the following constraints to represent the space of spanning trees. We assume $|V| = n$ and use $x = (x_e)_{e \in E}$ as the indicator

vector for the selected spanning tree, and use $E[L]$ to represent the set of edges with both endpoints in the vertex set L.

$$\sum_{e\in E} x_e = n-1 \tag{5a}$$

$$\sum_{e\in E[L]} x_e \le |L|-1, \quad \forall L \subseteq V \tag{5b}$$

$$x_e \in \{0,1\}, \quad \forall e \in E. \tag{5c}$$

The inner level problem of (4) is to solve the minimum spanning tree problem over the subgraph induced by the edge set $T \cup S$. Since the goal is to dualize the inner problem, we choose a compact linear formulation from [41] as follows,

$$\min \quad \sum_{e\in E} w_e y_e \tag{6a}$$

$$\sum_{\{j:\{i,j\}\in E\}} f_{ij}^l - \sum_{\{j:\{i,j\}\in E\}} f_{ji}^l = b_i^l \quad \forall i \in V,\, l \in V\setminus\{k\} \tag{6b}$$

$$y_e \ge f_{ij}^l + f_{ji}^{l'} \qquad \forall e=\{i,j\}\in E,\, l \ne l' \in V\setminus\{k\} \tag{6c}$$

$$\sum_{e\in E} y_e = n-1 \tag{6d}$$

$$f_{ij}^l, f_{ji}^l \ge 0 \qquad \forall \{i,j\}\in E,\, l \in V\setminus\{k\} \tag{6e}$$

$$y_e \ge 0 \qquad \forall e \in E, \tag{6f}$$

where $(y_e)_{e\in E}$ is the indicator vector of the spanning tree to be selected. In this setting, a root vertex k is arbitrarily selected from V, and a commodity l is defined for every non-root vertex $l \in V\setminus\{k\}$. Then, one unit of each commodity l emanating from the root vertex k must be delivered to the vertex l. Thus, the parameter b_i^l takes values of 1, -1, or 0 whenever $i=k$, $i=l$, or $i \in V\setminus\{k,l\}$, respectively. Finally, we use the following linking constraints to combine formulations (5) and (6),

$$y_e \le x_e + \mathbb{1}_S(e) \quad \forall e \in E, \tag{7}$$

where $\mathbb{1}_S$ is the indicator function of the input edge set S. Thus, an edge e can be used in the inner problem if either it has been selected as a part of the spanning tree by x_e or if it belongs to the input set S. Then, dualizing the inner problem produces a one-level mixed-integer problem (MIP). By Theorem 4, S is a critical nucleus if and only if this MIP formulation obtains an upper bound that is less than r. △

Example 2 (Shortest Path Interdiction, Cont.)**.** In this setting, the outer problem of (4) needs to impose a solution space of paths, and the inner problem is to solve for a shortest path on the subgraph restricted on $T \cup S$. Thus, we can use the following bilevel formulation.

$$\begin{aligned}
\max_{x\in\mathcal{X}} \quad \min \quad & \sum_{e\in E} w_e y_e \\
& \sum_{e\in\delta^+(i)} y_e - \sum_{e\in\delta^-(i)} y_e = b_i, \quad \forall i \in V \\
& y_e \le x_e + \mathbb{1}_S(e), \quad \forall e \in E \\
& y_e \ge 0, \quad \forall e \in E,
\end{aligned}$$

where b_i takes values 1, -1, and 0 whenever $i = s$, $i = t$, and $i \in V \setminus \{s, t\}$, and $\mathcal{X}$ is the space of indicator vectors of paths as follows,

$$\begin{aligned} \sum_{e \in \delta^+(i)} x_e - \sum_{e \in \delta^-(i)} x_e &= b_i, \quad \forall i \in V \\ \sum_{e \in \delta(i)} x_e &\leq 2, \quad \forall i \in V \setminus \{s, t\} \\ x_e &\in \{0, 1\}, \quad \forall e \in E. \end{aligned}$$

We use $\delta(i)$ to denote the set of all adjacent edges of i and make the variables integers to prevent the inner problem from potentially using fractional edges. Together with the flow balance requirements on s and t, the second constraint set eliminates all the potential cycles, which is essential in this outer problem for finding a maximum s-t path. In contrast, these constraints are relaxed in the inner problem since the minimization ensures that the optimal solutions would only indicate simple paths. Again, we can dualize the inner problem to produce a one-level MIP serving as an exact verifier for the $\mathcal{P}_\emptyset$-structures in this specific problem. △

Example 3 (Vertex Cover Interdiction, Cont.)**.** Let $N(v)$ and $N[v]$ denote the open and closed neighbors of a vertex $v \in V$ in the given graph G, i.e., $N[v] = N(v) \cup \{v\}$, a minimal vertex cover can be exactly characterized by the following.

Proposition 4. *Given $G = (V, E)$, a vertex cover $T \subseteq V$ is minimal if and only if*

$$|T \cap N[v]| \leq |N(v)|$$

for every $v \in V$.

Proof. It is known that the complement of a vertex cover is an independent set. Thus, T is a minimal vertex cover if and only if its complement $\bar{T}$ is a maximal independent set. According to [40], an independent set $\bar{T}$ is maximal if and only if $N[v] \cap \bar{T} \neq \emptyset$, which completes the proof. □

Using this characterization, we can use the following constraints to describe the space of minimal vertex covers and denote it as $\mathcal{X}$.

$$x_v + x_u \geq 1, \quad \forall (v, u) \in E \tag{8a}$$

$$\sum_{u \in N[v]} x_u \leq |N(v)|, \quad \forall v \in V \tag{8b}$$

$$x_v \in \{0, 1\}, \quad \forall v \in V. \tag{8c}$$

Accordingly, we can set up the following bilevel formulation.

$$\max_{x \in \mathcal{X}} \ \min_{y \in \mathcal{X}} \sum_{v \in V} w_v y_v \tag{9a}$$

$$y_v \leq x_v + \mathbb{1}_S(v), \quad \forall v \in V. \tag{9b}$$

This verifier is a bilevel integer linear program. We can use various bilevel branch-and-bound methods (e.g., [38, 65]) to solve this problem exactly. △

In general, suppose the solution space $m(\Omega)$ can be represented exactly by a solution space $\mathcal{X}$, then Formulation (4) can always be rewritten in the form of (9) using the ground set Δ in place of V, the ground set for the vertex cover example.

In all three examples, we have developed mathematical formulations to verify $\mathcal{P}_\emptyset$-structures. However, these verifiers are nowhere near practical since they need to be solved multiple times to separate new constraints. In fact, it is easy to see that solving some of these separation problems can be as challenging as solving the original problem.

To overcome such inefficiency, in the following subsection, we shift our focus to the development of partial verifiers that are allowed only to separate a subset of $\hat{\Omega}_{\mathcal{P}_\emptyset}$. To incorporate partial verifiers in Algorithm 1, we replace the MPS formulation with the MCS formulation in the algorithm, then generate supervalid inequalities using any given partial verifier. This variant algorithm searches for an optimal solution from some relaxation of Π_0, then compare this solution with the optimal solution obtained in Π_1. Hence, the overall correctness is still guaranteed.

4.3 A Partial Verifier for $\mathcal{P}_\emptyset$

Based on Theorem 3, in this section, we will provide two properties about the subsets of the ground set. We will prove that a structure $S \subseteq \Delta$ that holds both properties is guaranteed to be a $\mathcal{P}_\emptyset$-structure. This will lead to a more efficient procedure to identify $\mathcal{P}_\emptyset$-structures. The first concept is called *partial nuclei*.

Definition 8 (Partial Nucleus)**.** A set S is a partial nucleus of degree k, or a k-nucleus, if all elements in $m(\Omega)[S,k]$ are critical, i.e., $m(\Omega)[S,k] \subseteq \hat{\Omega}$. The set of k-nuclei is denoted by $\hat{\Omega}^c_k$.

Recall $m(\Omega)[S,k] = \{T \in m(\Omega) \mid |T \cap S| \geq |S| - k\}$ by Definition 6 for some $k \in \mathbb{N}$. In particular, if S is a 0-nucleus, then all minimal structures that contain S are critical, i.e., $m(\Omega)[S] \subseteq \hat{\Omega}$. Also, S is a k-nucleus for some $k \geq |S|$ if all minimal structures $m(\Omega)$ are critical, in which case all feasible solutions are in Π_1. Again, consider the edges set $S = \{e_1, e_2\}$ in Figure 2. Since no path in this example contains S, we trivially have $m(\Omega)[S,0] = \emptyset$, which implies that S is a 0-nucleus. On the other hand, because the non-critical path $\{e_1, e_3, e_5\}$ intersects S more than $|S| - 1 = 1$ edges, S is not a 1-nucleus. By this definition, we clearly have the following.

Proposition 5. *$\hat{\Omega}^c_k$ is an upper set and is decreasing on k. That is, for $k, k' \in \mathbb{N}$ such that $k \leq k'$, we have $\hat{\Omega}^c_k \supseteq \hat{\Omega}^c_{k'}$.*

Proof. For the first claim, for some $S \in \hat{\Omega}^c_k$ and any $S' \supseteq S$, we have $m(\Omega)[S',k] \subseteq m(\Omega)[S,k] \subseteq \hat{\Omega}$ by Proposition 3. Thus, $S' \in \hat{\Omega}^c_k$. For the second statement, $S \in \hat{\Omega}^c_{k'}$ means $m(\Omega)[S,k'] \subseteq \hat{\Omega}$. By Proposition 3, $k \leq k'$ implies

$$m(\Omega)[S,k] \subseteq m(\Omega)[S,k'] \subseteq \hat{\Omega}.$$

That is, $S \in \hat{\Omega}^c_k$. □

There are two reasons why the membership of a k-nucleus is much easier to verify than the critical nuclei. First, minimal elements $m(\Omega)$ are follower's structures and thus characterized by the follower's problem, whereas the elements in $m(\uparrow \Omega[S])$ can be arbitrary supersets of the follower's structures. Second, we can check whether any upper bound of the following

$$\max_{T \in m(\Omega)[S,k]} w(T) \tag{10}$$

is less than the interdiction target r to verify a k-nucleus instead of solving a bilevel problem, as it is required to verify a critical nucleus. Unfortunately, being a k-nucleus does not suffice to

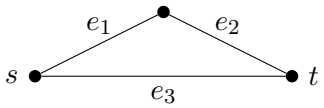
guarantee $S \in \hat{\Omega}_{\mathcal{P}_\emptyset}$. The following proposition investigates the relationship between k-nuclei and critical nuclei.

Proposition 6. $\hat{\Omega}_{\mathcal{P}_\emptyset} \subseteq \hat{\Omega}_0^c$. *Furthermore, there exist binary interdiction games where this inclusion is strict.*

Proof. Notice that when $k = 0$, we have $m(\Omega)[S, k] = m(\Omega)[S]$.

To prove the first statement, for every $S \in \hat{\Omega}_{\mathcal{P}_\emptyset}$, consider the set $m(\Omega)[S]$. This set is either empty or not. For the former, we have $\emptyset \subseteq \hat{\Omega}$, which makes S a 0-nucleus. For the latter, for every $T \in m(\Omega)[S]$, we have $T \in (\uparrow \Omega)[S]$. Because S is a critical nucleus, by definition, T must contain some critical $T' \subseteq T$. But, T' must be taken as T; otherwise, it will contradict the minimality of T in Ω. This shows $m(\Omega)[S] \subseteq \hat{\Omega}$, i.e., $S \in \hat{\Omega}_0^c$.

For the second statement, we construct the following instance of the shortest path interdiction game as a counterexample.



We assume $\hat{\Omega} = \{\{e_1, e_2\}\}$. Clearly, $\{e_1\}$ is a 0-nucleus since $m(\Omega)[\{e_1\}] = \{\{e_1, e_2\}\} = \hat{\Omega}$. However, $\{e_1\}$ is not a critical nucleus since taking T as the path $\{e_3\}$, the union $T \cup \{e_1\}$ does not contain any critical paths. □

By this proposition, Formulation (10) is not a valid partial verifier as it may accept "false positive" structures. To resolve this issue, we use the following definition.

Definition 9 (Regenerability)**.** A set S is k-regenerable in Ω, if for all $T \in \Omega$, there is a $T' \subseteq T \cup S$ such that $T' \in m(\Omega)[S, k]$. The family of k-regenerable sets is denoted by $\hat{\Omega}_k^r$.

Intuitively, a set S is k-regenerable if combining it with any structure $T \in \Omega$ by union will contain a minimal structure $T' \in m(\Omega)$ that intersects S with at least $|S| - k$ elements. In particular, 0-regenerable implies $S \cup T$ always regenerates some minimal structure T' that contains S as a subset, which is quite a strong condition. On the other hand, every set S is k-regenerable for all $k \geq |S|$ since $S \cup T$ must contain some minimal structure (either T or some minimal structure contained in T). This implies, in contrast to the k-nuclei, $\hat{\Omega}_k^r$ is increasing on k. To be exact, we have the following result for $\hat{\Omega}_k^r$.

Proposition 7. $\hat{\Omega}_k^r$ *is increasing on* k*. Moreover, for every* $S \in \hat{\Omega}_k^r$ *and* $S' \supseteq S$*, we have* $S' \in \hat{\Omega}_{k+|S'|-|S|}^r$.

Proof. To prove the first, we take any $S \in \hat{\Omega}_k^r$. By definition, for every $T \in \Omega$, there is some $T' \subseteq T \cup S$ such that $T' \in m(\Omega)[S, k]$. By Proposition 3, we have $m(\Omega)[S, k] \subseteq m(\Omega)[S, k']$ for $k' \geq k$. Thus, the same T' also exists in $m(\Omega)[S, k']$. By definition, $S \in \hat{\Omega}_{k'}^r$. For the second statement, for each $T \cup S'$, we can take the same T' that is regenerated by $T \cup S$, which intersects T' with at least

$$|S| - k = |S'| - (k + |S'| - |S|)$$

elements, which completes the proof. □

Combining partial nuclei with regenerability, we get the following verification theorem.

Theorem 5 (Partial Verification Theorem)**.** $\hat{\Omega}_k^c \cap \hat{\Omega}_k^r \subseteq \hat{\Omega}_{\mathcal{P}_\emptyset}$ *for all* $k \in \mathbb{N}$.

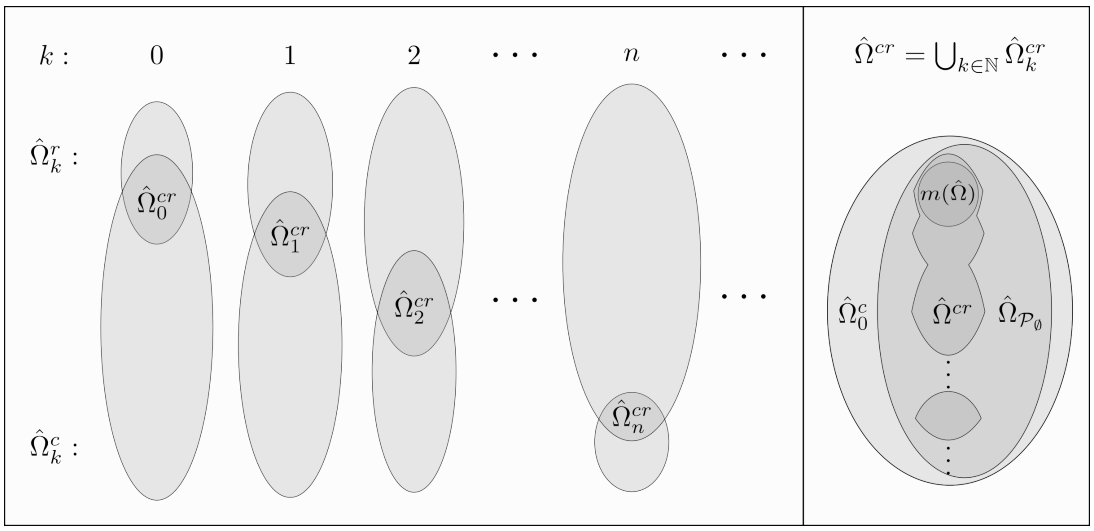


Figure 6: Hierarchy of Partial Nuclei $\hat{\Omega}_k^c$ and Regenerable Sets $\hat{\Omega}_k^r$. The left figure shows that $\hat{\Omega}_k^c$ and $\hat{\Omega}_k^r$ are decreasing and increasing on k, respectively. In the right figure, their intersections $\hat{\Omega}_k^{cr}$'s are combined into the set $\hat{\Omega}^{cr}$, which is sandwiched between $m(\hat{\Omega})$ and $\hat{\Omega}_{\mathcal{P}_\emptyset}$. In addition, the set of 0-nuclei contains all the critical nuclei $\hat{\Omega}_{\mathcal{P}_\emptyset}$ as a subset.

Proof. Given a set $S \in \hat{\Omega}_k^c \cap \hat{\Omega}_k^r$ for some k, suppose $S \notin \hat{\Omega}_{\mathcal{P}_\emptyset}$, i.e., S is not a cut of the corresponding $\hat{\Pi}_0$. Then, there is some feasible strategy $U \in \hat{\Pi}_0$ that does not intersect S. Because $\mathcal{P}_\emptyset(U) = 0$, by the definition of null property, some structure $T \in \Omega$ is not blocked by U. Thus, we have $(S \cup T) \cap U = \emptyset$. Then, S is k-regenerable in Ω implies there is some $T' \subseteq S \cup T$ such that $T' \in m(\Omega)[S, k]$. Moreover, S is a k-nucleus implies such T' must be critical. Note $U \cap T' = \emptyset$ by construction, which contradicts the feasibility of U. So, we are done. □

Corollary 4. *For any $k \in \mathbb{N}$, denote $\hat{\Omega}_k^{cr} := \hat{\Omega}_k^c \cap \hat{\Omega}_k^r$ and define $\hat{\Omega}^{cr} := \bigcup_{k\in\mathbb{N}} \hat{\Omega}_k^{cr}$. Then, $\hat{\Omega}^{cr} \subseteq \hat{\Omega}_{\mathcal{P}_\emptyset}$.*

We omit the proof as it is a trivial consequence of Theorem 5. In particular, this corollary provides an upper bound of $\hat{\Omega}^{cr}$. On the other hand, the following proposition provides a lower bound, which can be verified directly from the definitions.

Proposition 8. *Every critical structure $T \in m(\hat{\Omega})$ is a 0-nucleus and is 0-regenerable in Ω. In particular, $m(\hat{\Omega}) \subseteq \hat{\Omega}_0^{cr} \subseteq \hat{\Omega}^{cr}$.*

The relationship between all these sets forms a hierarchy map illustrated in Figure 6. From a practical perspective, Theorem 5 enables us to split the verification procedure into two parts. Given a set S, we only need to check: (i) whether it is a k-nucleus using Formulation (10) or any equivalent algorithms, (ii) whether it is k-regenerable for certain k between 0 and $|S|$. For the latter, we can often circumvent the verification step by starting with structures that are known to be k-regenerable. We provide several examples in the next subsection to illustrate these ideas.

4.4 Examples for the Partial Verifier

Given a set S, we can use the following general formulation to determine the smallest k required for S to be k-regenerable.

$$k = |S| - \min_{T\in m(\Omega)} \max_{\substack{T'\subseteq S\cup T \\ T'\in m(\Omega)}} |S \cap T'|.$$

However, this is again a bilevel optimization similar to the exact verifier (4). In practice, we can often skip this verification by identifying k-regenerable structures directly, and then applying partial nucleus verifier (10) on these regenerable structures to check whether S is a $\mathcal{P}_\emptyset$-structure.

Example 1 (Spanning Tree Interdiction, Cont.)**.** In this case, every edge set S that does not contain a cycle is 0-regenerable due to the edge exchange property of spanning trees. Therefore, we can start with a critical spanning tree $T \in \hat{\Omega}$, then verify its proper subsets using any implementation of the partial nucleus verifier (10). This verifier computes the maximum weight of the spanning trees that contain S. A variant of Kruskal's algorithm has been implemented in [59]. In the same paper, the authors demonstrated that, with this $\mathcal{P}$-structure verifier for separating Constraint (3b), the resulting MPS formulation works more efficiently than the MCS formulation. △

Example 2 (Shortest Path Interdiction, Cont.)**.** In contrast to the minimum spanning tree case, not every subset of a s-t path is 0-regenerable. To identify the regenerable sets in this problem, we use the following definitions.

Definition 10 (Path Decomposition & Skeletons)**.** Given a graph $G = (V, E)$ and a source-terminal pair $s, t \in V$, a *path decomposition* is a sequence of vertex subsets $\{X_i\}_{i=0}^n$ with three properties:

- $s \in X_0$ and $t \in X_n$;
- for every $\{u, v\} \in E$, there exists some i such that $u, v \in X_i$;
- for every $i \le j \le k$, $X_i \cap X_k \subseteq X_j$.

For each X_i, the *source set* S_i and *terminal set* T_i are defined as

$$S_i = \begin{cases} \{s\}, & i = 0 \\ X_i \cap X_{i-1}, & \text{otherwise.} \end{cases} \qquad T_i = \begin{cases} \{t\}, & i = n \\ X_i \cap X_{i+1}, & \text{otherwise.} \end{cases}$$

For a set of vertices X, we define $E(X) := \{\{u, v\} \in E \mid u, v \in X\}$. Then, for any part X_i such that $S_i \cap T_i = \emptyset$, a *skeleton* K of X_i is a minimal edge set that preserves the connectivity between S_i and T_i, i.e., for every part $(s, t) \in S_i \times T_i$ that is connected through edges in $E(X_i) \setminus (E(S_i) \cup E(T_i))$, there is a path in K that connects them, and K is a minimal edge set that achieves this. Let $\mathcal{P}_K$ be all such paths in K, then we define

$$\lambda(K) := \min_{P \in \mathcal{P}_K} |P|,$$

i.e., the length of the shortest path (in terms of the number of edges) in K that connects S_i to T_i.

This definition is a variant of the path decomposition in [48], and we provide an example in Figure 7 for illustration. The intuition is that a path decomposition reduces the given graph G into a simple s-t path on the aggregated level. In particular, for every $i \le j \le k$, a path from some $u \in X_i$ to some $v \in X_k$ must take some subpath in X_j with a source vertex in S_i and a terminal vertex in T_i. This decomposition is related to regenerable sets by the following proposition.

Proposition 9. *In the shortest path interdiction, for an arbitrary path decomposition, every skeleton K induces a $(|K| - \lambda(K))$-regenerable set.*

Proof. For every s-t path P and a skeleton K with respect to a part X_i under some path decomposition, the subpath of P that passes through some X_i such that $S_i \cap T_i = \emptyset$ must connect some vertex $u \in S_i$ to some vertex in $v \in T_i$. Swapping this subpath with the unique (u, v)-path in K produces a new path from s to t. Moreover, the (u, v)-path in K has a length that is greater or equal to $\lambda(K)$. Hence, the new path intersects with K in at least $\lambda(K) = |K| - (|K| - \lambda(K))$ edges, which proves the claim. □

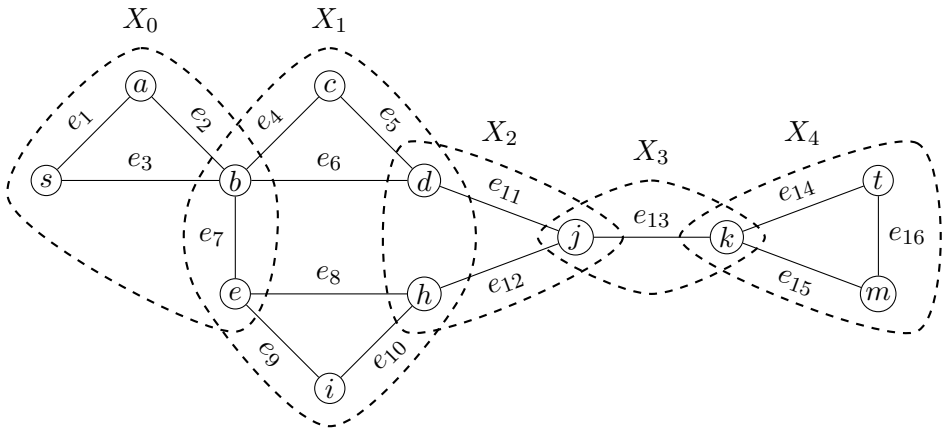


Figure 7: Path decomposition and skeletons. All the vertices are covered by the path decomposition $\{X_i\}_{i=0}^4$. For X_1, the source set S_1 and terminal set T_1 are $\{b, e\}$ and $\{d, h\}$, respectively. $K_1 = \{e_6, e_8\}$, $K_2 = \{e_4, e_5, e_9, e_{10}\}$ are both skeletons in X_1 since both are minimal edge sets that preserve all the connectivities between S_1 and T_1. By definition, we have $\lambda(K_1) = 1$ and $\lambda(K_2) = 2$. In contrast, $\{e_6\}$ and $\{e_4, e_5, e_6, e_8\}$ are not skeletons as the former breaks the connectivity between e and h while the latter is not minimal. Another example, for X_0, S_0 is $\{s\}$ and $T_0 = \{b, e\}$. Both $K_3 = \{e_1, e_2\}$ and $K_4 = \{e_3\}$ are skeletons with $\lambda(K_3) = 2$ and $\lambda(K_4) = 1$, while $\{e_3, e_7\}$ is not since it contains edges in T_0, thus is not minimal.

For instance, in Figure 7, $\{e_1, e_2\}, \{e_3\}, \{e_{13}\}$ are all 0-regenerable; $\{e_6, e_8\}$ and $\{e_{11}, e_{12}\}$ are both 1-regenerable; $\{e_4, e_5, e_8\}$ and $\{e_4, e_5, e_9, e_{10}\}$ are both 2-regenerable. Notice that some regenerable sets (e.g., $\{e_{11}, e_{12}\}$ or $\{e_{13}\}$) are uninteresting since they are k-nucleus if and only if all paths are critical. In general, it is easy to see that every s-t cut S is $(|S| - 1)$-regenerable. Yet, they are all trivial regenerable sets for the same reason.

Using this definition, we can randomly construct several path decompositions in this problem, then generate non-trivial skeletons from some parts. For each skeleton K, we can run Formulation (10) or any equivalent variant of the longest path algorithm to intersect K with at least $|K| - \lambda(K)$ edges. Suppose the resulting length is less than the predefined r, we separate a constraint of (3b) that is associated with K. We also note that these algorithms can be terminated early once any upper bound is found to be less than r. We provide the corresponding computational experiments in Appendix 6, where this supervalid inequality approach based on the path decomposition is compared with the standard MCS formulation (2) tailored for the shortest path interdiction problem. △

Example 3 (Vertex Cover Interdiction, Cont.)**.** Similar to the shortest path case, not every subset of a vertex cover is 0-regenerable. For instance, $T_1 = \{v_1, v_2, v_4, v_5\}$ and $T_2 = \{v_1, v_2, v_3, v_6\}$ are both minimal vertex covers in Figure 1. Take $\{v_5\} \subseteq T_1$, but $T_2 \cup \{v_5\}$ does not contain any minimal vertex cover that contains $\{v_5\}$ since all the neighbors of v_4 (i.e., elements in T_2) have to be included in a minimal vertex cover that excludes v_4. The following definition and lemma identify a certain type of vertex that is 0-regenerable. Recall that, for a vertex $v \in V$, we use $N(v)$ and $N[v]$ for the open and closed neighbors of v in the graph.

Definition 11 (Sociable Vertices)**.** In a simple graph $G = (V, E)$, a vertex $v \in V$ is called *sociable* if there exists $u \in N(v)$ such that $N[u] \subseteq N[v]$.

Lemma 1. *Every singleton $\{v\}$ for some sociable vertex v is 0-regenerable.*

Proof. Pick any minimal vertex cover $T \in m(\Omega)$, either $v \in T$ or not. The former is trivial as we can pick $T' = T \subseteq T \cup \{v\}$. Suppose otherwise, we have $N(v) \subseteq T$ to cover all the edges between v

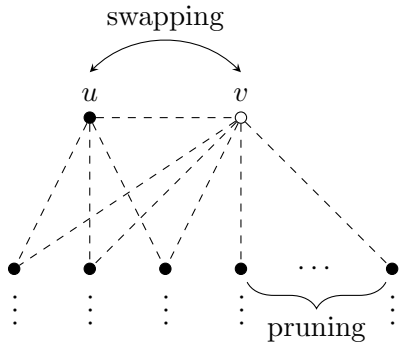


Figure 8: Swapping and pruning operations on the sociable vertex v.

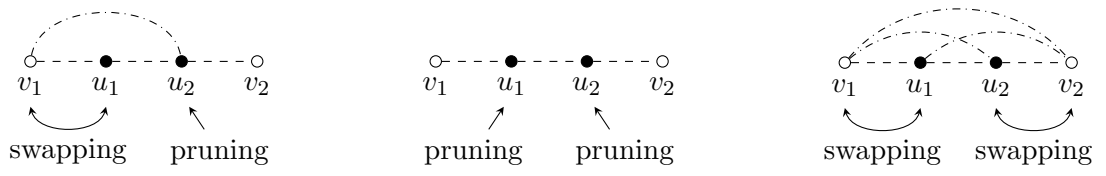
and its neighbors. Take one vertex $u \in N(v)$ such that $N[u] \subseteq N[v]$, we swap u with v to construct $T_0 = T \cup \{v\} \setminus \{u\}$. We will show that (i) T_0 is a vertex cover and (ii) it contains some minimal vertex cover T' that also contains v. For (i), notice that by this *swapping* operation (see Figure 8), only the coverage of edges $\{v\} \times N(v)$ and $\{u\} \times N(u)$ has been changed. We partition these edges into three parts,

1. (u, v);
2. $\{v\} \times (N(v) \setminus \{u\})$;
3. $\{u\} \times (N(u) \setminus \{v\})$.

This is clearly a partition. For 1, after swapping u and v, the edge $\{u, v\}$ is still covered; for 2, note that $N(v) \setminus \{u\} \subseteq T$ since $v \notin T$, and swapping u and v would not change this containment, i.e., $N(v) \setminus \{u\} \subseteq T_0$. Thus, all the edges in Case 2 are covered by T_0. Case 3 is trivial as $N[u] \subseteq N[v]$. To prove (ii), we use Proposition 4. That is, we will construct some T' that satisfies $v \in T' \subseteq T_0$ such that, for each $v \in V$, there exists some $v' \in N[v]$ does not belong to T'. Because only the vertices in $N[v] \cup N[u]$ have been affected by the swapping operation, we partition these vertices as follows,

1. $\{u, v\}$;
2. $N(u) \setminus \{v\}$;
3. $N(v) \setminus N[u]$, which is empty if $N[v] = N[u]$.

This is a valid partition (that may contain empty sets) since $N[u] \subseteq N[v]$. For Case 1, both u and v satisfy the characterization as $u \notin T_0$. All the vertices in Case 2 also satisfy the characterization as they are neighbors of u. For Case 3, suppose $N(v) \setminus N[u]$ is nonempty (otherwise we are done), we can sort $N(v) \setminus N[u]$ in an arbitrary order. Then, we perform a *pruning* operation on $N(v) \setminus N[u]$ sequentially. Take each element v' from this set in order, either $N[v'] \subseteq T_0$ or not. For the latter, we do nothing and go to the next element in the set. For the former, we remove v' from T_0. Notice that $T_0 \setminus \{v'\}$ is still a vertex cover, since removing v' only affects its neighbors. But, the reason to remove v' is that all its neighbors are covered in T_0, which implies all the edges $\{v'\} \times N(v')$ are still covered by $T_0 \setminus \{v'\}$. Clearly, after scanning all the elements in $N(v) \setminus N[u]$, we are left with a set T' where every vertex satisfies the characterization, and T' is preserved all along as a vertex cover. Moreover, since the only operation is removing vertices in T_0 other than v, we have $v \in T' \subseteq T_0$, which implies v is 0-regenerable. □



(a) Case 1 (b) Case 2 (c) Case 3

Figure 9: Three cases in Proposition 10. The dash-dotted edges exist due to $N[u_i] \subseteq N[v_i]$, which is implied by the corresponding swapping operations.

However, a 0-regenerable singleton T is often weak in its quality since the set $\Omega[T]$ could be quite large, which is unlikely to be contained in $\hat{\Omega}$. The following proposition shows that these sociable vertices can be merged to a larger 0-regenerable set as long as they keep a certain distance from each other.

Proposition 10. *A set of sociable vertices $U \subseteq V$ is 0-regenerable if it is also an independent set.*

Proof. For every minimal vertex cover T and every sociable vertex $v \in U$, the swapping and pruning operations, if performed, will only add v into T and remove some neighbors in $N(v)$ from T (including the swapped u). Thus, we only need to show that, after sequentially performing the two operations on multiple sociable vertices in U, one of the two endpoints of every edge is still covered by the resulting T'. Since all the operations within each neighbor are safe by Lemma 1, the only possibility to cause some edge to be uncovered is that, for some $v_1, v_2 \in U$ such that $v_1 \neq v_2$, one vertex $u_1 \in N(v_1)$ removed from T along with some $u_2 \in N(v_2)$ removed from T form an edge (u_1, u_2). More specifically, we have the following three possible cases (see Figure 9):

1. without loss of generality, u_1 and u_2 are removed from T due to swapping and pruning operations, respectively;
2. both u_1 and u_2 are removed from T due to pruning operations;
3. both u_1 and u_2 are removed from T due to swapping operations.

We will show all three cases are impossible to occur. Consider we scan each $v \in U \setminus T$ in a given order and perform the corresponding swapping and pruning operations in each iteration. First, these operations will not remove any vertex in $U \cap T$ from T, because these operations only affect the neighbors of v while the minimum distance between vertices in U is at least two by the assumption that U is an independent set.

Then, for Case 1, either the swapping operation on v_1 is performed before or after the pruning operation on $N(v_2)$. For the former, after the swap, u_1, as a neighbor of u_2, has been removed from T. Thus, u_2 will not be pruned afterward as not all its neighbors are selected. For the latter, since v_1 is sociable and u_1 will be swapped with v_1, we have $N[u_1] \subseteq N[v_1]$ by definition. Clearly, the edge (u_1, u_2) is adjacent to u_1, thus we also have $u_2 \in N[v_1]$, i.e., $v_1 \in N(u_2)$. Before the swap, we further have $v_1 \notin T$. Thus, u_2 will not be removed from T by the pruning process as not all the neighbors of u_2 are selected (v_1 is not). This completes Case 1.

For Case 2, no matter which pruning operation happens first (say u_1), the other pruning operation will not be carried out since its neighbor u_1 is not selected in the constructed vertex cover.

For Case 3, since u_1 and u_2 in T will be swapped with v_1 and v_2, respectively. We have $N[u_i] \subseteq N[v_i]$ for $i = 1, 2$. In particular, $u_1 \in N[v_2]$ and $u_2 \in N[v_1]$. Moreover, v_1 is also a neighbor of u_2, so we have

$$v_1 \in N[u_2] \subseteq N[v_2],$$

which contradicts the minimum distance between v_1 and v_2.

Therefore, all three cases are impossible to occur, which implies all the procedures of scanning $v \in U \setminus T$ and performing the corresponding swapping and pruning operations will never lead to any conflicts. Then, by Lemma 1, after the entire process, the resulting T' is still a vertex cover that satisfies the characterization of Proposition 4, thus is a minimal vertex cover that satisfies $U \subseteq T' \subseteq (U \cup T)$. This completes the proof. □

Accordingly, in this problem, we can first identify all (or part of) the sociable vertices U, then pre-compute their pairwise distance. Then, for every $U' \subseteq U$ that satisfies the distance requirement, we can run the following 0-nucleus verifier

$$\begin{aligned} \max_{x \in \mathcal{X}} \ & \sum_{v \in V} w_v x_v \\ & \text{(8a)–(8c)}, \\ & x_v = 1, \quad \forall v \in U'. \end{aligned}$$

Suppose the optimal value or any upper bound is less than the predefined target r, then U' is a $\mathcal{P}_\emptyset$-structure. Compared to the bilevel integer program (9), this verifier is easier to solve. △

4.5 General $\mathcal{P}_\emptyset$-Structure Separation Guidelines

In general, for a specific binary interdiction game, we can separate the $\mathcal{P}_\emptyset$-structures using the following general guidelines. As long as (2) is sufficient for producing an optimal solution, these supervalid inequalities could be incorporated to speed up the overall solution procedure.

1. Identify the set of follower's structures Ω.
2. Identify a regular bipartition property $\mathcal{P}$ such that the optimal value $z(\Pi_1)$ can be efficiently determined.
3. Identify a class of regenerable sets in the corresponding problem.
4. Solve the MPS formulation (3) using Algorithm 1.
5. At each iteration of the separation step, heuristically select some regenerable set S, then run the corresponding k-nucleus verifier (10) on S. If an upper bound of (10) is obtained to be less than the interdiction target r, then we add the constraint $\sum_{a \in S} x_a \geq 1$ into the MPS formulation.
6. The algorithm terminates when the MPS formulation solves optimally, or its lower bound is no less than $z(\Pi_1)$.

All the previous examples can set up a separation subroutine based on these guidelines with deliberately designed implementations. For instance, in the minimum spanning tree interdiction [60], the authors developed an enumeration tree for the regenerable sets and implemented a binary search on each branch of the enumeration tree to separate the $\mathcal{P}_\emptyset$-structures. Though the regenerable set selection and k-nucleus verifier implementation for different binary interdiction games could be vastly different, a special class of structures, called the *greedoids*, shares similar features as the spanning trees in the $\mathcal{P}_\emptyset$-structure separation procedure.

5 Greedoid Interdiction Games

Previous examples imply that an efficient implementation of the partial verifier requires (i) a set of easy-to-identify regenerable sets and (ii) an efficient algorithm to obtain a tight upper bound for (10). In this section, we will show that, for the special class of problems called the *greedoid interdiction games*, both requirements are satisfied under mild assumptions.

5.1 Preliminaries on Greedoids

A greedoid is an abstract set system that connects algorithms, combinatoric optimizations, and classical analysis of mathematical structures [36]. Intuitively, greedoids extract the fundamental properties of problems where greedy algorithms can produce optimal solutions [11]. It also generalizes many other abstract set systems, such as interval greedoids, antimatroids, and matroids. Many applications and research fields are associated with these set systems, including graph theory [57], electric network theory [47], assignment problems [25], game theory [1], semi-Markov process [27], poset analysis [6], and many more. We begin with the following definitions.

Definition 12 (Greedoid/Basis/Rank/Closure/Greedoid Optimization [36, p. 10, 44, 158, 160])**.** A set system $(\Delta, \widetilde{\Omega})$ with $\widetilde{\Omega} \subseteq 2^{\Delta}$ is called a *greedoid* if it satisfies the following two properties,

- *Accessibility*: for every $T \in \widetilde{\Omega}$ there exists $a \in T$ such that $T \setminus \{a\} \in \widetilde{\Omega}$;
- *Exchange property*: for every $T, T' \in \widetilde{\Omega}$ with $|T| < |T'|$, there is some $a \in T' \setminus T$ such that $T \cup \{a\} \in \widetilde{\Omega}$.

Given a greedoid $(\Delta, \widetilde{\Omega})$, all the maximal elements of $\widetilde{\Omega}$ form the set of *bases* denoted by Ω. The *rank function* $r : 2^{\Delta} \to \mathbb{N}$ and *closure operator* $\sigma : 2^{\Delta} \to 2^{\Delta}$ are defined as

$$r(S) = \max\left\{|T| \mid T \subseteq S,\ T \in \widetilde{\Omega}\right\},$$
$$\sigma(S) = \{a \in \Delta \mid r(S \cup \{a\}) = r(S)\}.$$

A set S is *closed* if $\sigma(S) = S$. We say $\widetilde{\Omega}$ has the *strong exchange property* ([36, p. 27]) if for every $S \subseteq T \in \Omega$ with $S \in \widetilde{\Omega}$ and every $a \in \Delta \setminus T$ with $S \cup \{a\} \in \widetilde{\Omega}$, there exists some $b \in T \setminus S$ such that both $S \cup \{b\}$ and $(T \setminus \{b\}) \cup \{a\}$ are in $\widetilde{\Omega}$. The corresponding *greedoid optimization* is defined as

$$\max_{T \in \Omega}\ w(T),$$

where Ω is the set of bases of $\widetilde{\Omega}$ and $w(\cdot)$ is any weight function. We say $w(\cdot)$ is linear if $w(T) = \sum_{a \in T} w(a)$ for some weight assignment of the ground set $w : \Delta \to \mathbb{R}$.

By this definition, it is easy to verify that all bases of a greedoid $\widetilde{\Omega}$ have the same size, which is also called the rank of $\widetilde{\Omega}$. Because of the exchange property, every non-basis element in a greedoid $\widetilde{\Omega}$ can be extended gradually to a basis. That is, for every $T \in \widetilde{\Omega}$, define

$$\Gamma(T) := \{a \in \Delta \setminus T \mid T \cup \{a\} \in \widetilde{\Omega}\}.$$

Then, $\Gamma(T)$ is empty if and only if T is a greedoid basis. This observation enables a natural greedy algorithm (i.e., Algorithm 2) for solving the greedoid optimization problem. The following proposition gives one of the main results regarding the greedy algorithm on greedoids.

Algorithm 2: Greedy Algorithm for Greedoids [36, p. 160]

Data: input $\Delta, \widetilde{\Omega}, w(\cdot)$
$T \leftarrow \emptyset$ // start with the minimal structure in $\widetilde{\Omega}$
while $\Gamma(T) \neq \emptyset$ **do**
  $a^\star \leftarrow \max_{a \in \Gamma(T)} w(a)$ // select a greedy choice
  $T \leftarrow T \cup \{a^\star\}$
end
return T

Proposition 11 ([36, p. 160])**.** *Let $(\Delta, \widetilde{\Omega})$ be a greedoid with Ω as its bases. Then the following are equivalent:*

- *For every linear objective function $w(\cdot)$, the greedy algorithm is optimal.*
- *$\downarrow \Omega$ is a matroid and every closed set in $(\Delta, \widetilde{\Omega})$ is also closed in $(\Delta, \downarrow \Omega)$.*
- *$(\Delta, \widetilde{\Omega})$ has the strong exchange property.*

5.2 Partial Verifier in Greedoid Interdiction Games

We can naturally associate a binary interdiction game with each greedoid structure $\widetilde{\Omega}$. We call this problem the *greedoid interdiction game* with respect to $\widetilde{\Omega}$ and is defined as follows.

Definition 13 (Greedoid Interdiction Games)**.** Given a greedoid $(\Delta, \widetilde{\Omega})$ with weights $\{w_a\}_{a \in \Delta}$ and an interdiction target r, the corresponding *greedoid interdiction* is a binary interdiction game where the set of minimal structures consists of all the bases in $\widetilde{\Omega}$ with the objective to interdict every structure $T \in \widetilde{\Omega}$ such that $w(T) = \sum_{a \in T} w_a < r$.

The following theorem shows that the regenerable sets can be effortlessly identified in greedoid interdiction games.

Theorem 6. *Given a greedoid $(\Delta, \widetilde{\Omega})$, every $S \in \widetilde{\Omega}$ is 0-regenerable in the associated greedoid interdiction game.*

Proof. Let Ω be the set of bases in $\widetilde{\Omega}$. By the exchange property of greedoid, for every feasible set $S \in \widetilde{\Omega}$ and every basis $T \in \Omega$, $S \cup T$ contains some $T' \in \Omega$ that satisfies $T' \supseteq S$, which implies S is 0-regenerable in the associated greedoid interdiction game. □

Therefore, for the binary interdiction game associated with a greedoid $\widetilde{\Omega}$, we can search for some proper subset $S \subseteq T \in m(\hat{\Omega})$ that satisfies $S \in \widetilde{\Omega}$, then run the 0-nucleus verifier (10) over S and accept it as a $\mathcal{P}_\emptyset$-structure if any upper bound of the verifier gets less than the predefined target value r. Moreover, according to the following theorem, suppose that the greedoid $\widetilde{\Omega}$ further satisfies the strong exchange property, then 0-nucleus verifier can be implemented as a greedy algorithm.

Theorem 7. *For the greedoid interdiction game associated with a greedoid $(\Delta, \widetilde{\Omega})$ that satisfies the strong exchange property, the corresponding 0-nucleus verifier (10) with a linear weight function $w(\cdot)$ adopts a greedy algorithm that solves for an optimal solution.*

Proof. Let Ω be the set of bases in the greedoid $\widetilde{\Omega}$. Then, the 0-nucleus verifier (10) for the associated greedoid interdiction game can be written as

$$\max_{T\in\Omega[S]} w(T). \tag{11}$$

By Proposition 11, the strong exchange property implies $\downarrow\Omega$ is a matroid and every closed set in $(\Delta,\widetilde{\Omega})$ is also closed in $(\Delta,\downarrow\Omega)$. Fixing partial solution S in the matroid $\downarrow\Omega$ is also called the contraction operation relative to S, which induces the contracted matroid $(\Delta\setminus S,(\downarrow\Omega)/S)$ with

$$(\downarrow\Omega)/S := \{T'\subseteq\Delta\setminus S \mid T'\cup T\in\downarrow\Omega \text{ for some } T\in\mathcal{B}(S)\},$$

where $\mathcal{B}(S)$ are the maximal elements in $\{T\subseteq S \mid T\in\widetilde{\Omega}\}$ (see [36, p. 15]). Thus, (11) is the greedoid optimization problem associated with this contracted matroid. Moreover, the closed sets in the contracted matroid are inherited from the closed sets in $\downarrow\Omega$. This means the contracted matroid satisfies the second condition in Proposition 11. Thus, the greedy algorithm is optimal to solve (11). □

These two theorems demonstrate that the $\mathcal{P}_\emptyset$-structure separation is often efficient for greedoid interdiction games.

6 Extension in Interdiction Games with a Limited Budget

Thus far, we have derived the existence and separation of supervalid inequalities solely for interdiction games formulated as (1) with a predefined interdiction target value r. However, many interdiction games in the literature are instead formulated in terms of a limited budget b, which can be generally described as follows.

$$z^\star := \max_{U\in\Pi:c(U)\le b} \min_{T\in\Omega\setminus\Omega_U} w(T). \tag{12}$$

This section explores the application of the proposed $\mathcal{P}$-structures in this type of interdiction game.

According to [59], regardless of whether the inner problem is linear or binary, (12) can be equivalently reformulated into Formulation (2) by setting $r=z^\star$. It is worth noting that the constraints in (2b) corresponding to $\hat{\Omega}_{z^\star}$ have been identified as the traditional supervalid inequalities in the literature [14, 32] for the budget version of interdiction games.

In implementation, the unknown value $z^\star$ along with the corresponding set of critical structures $\hat{\Omega}_{z^\star}$ can be iteratively refined using the incumbent objective value. Specifically, the cut generation method based on the following master and sub-problem pair can be used to solve the budget version.

$$\begin{aligned}\max_{x\in\{0,1\}^n:c(x)\le b}\quad & c(x)\\ \text{s.t.}\quad & \sum_{a\in T} x_a \ge 1 \quad \forall T\in\hat{\Omega}_t,\end{aligned} \qquad\qquad \min_{T\in\Omega\setminus\Omega_x}\quad w(T).$$

We initialize the master problem with $\hat{\Omega}_0=\emptyset$. Then, in each iteration t, we solve the master problem under $\hat{\Omega}_t$ to obtain a solution x, then we generate the optimal structure T from the subproblem under the current solution x to form $\hat{\Omega}_{t+1} := \hat{\Omega}_t\cup\{T\}$ for the next iteration. This procedure terminates when the master problem or the sub-problem becomes infeasible and then returns the incumbent as the optimal interdiction plan. In this sense, the budget version can be considered as a special case of the target version (2). Intuitively, this approach encourages to keep pushing the lower bound of the inner problem by changing the leader's strategies until either running out

all the budget or blocking all the follower's structures. Since each iteration of this implementation is essentially a feasibility test, the objective function can be arbitrary. We choose $\max_{x:c(x)\leq b} c(x)$ simply to speed up the algorithm by exploring only the maximal interdiction strategies. We note that similar ideas based on this sample-generation method have also been explored in several recent papers [23, 39].

As a consequence, the supervalid inequalities obtained from the proposed partial verifier can still be used in this cut generation procedure according to the following two propositions.

Proposition 12. *Given follower's critical structures $\hat{\Omega} \subseteq 2^{\Delta}$ and leader's strategies represented as the intersection $\Pi \cap \Pi'$, let $\mathcal{P}$ be a bipartition on Π with $\hat{\Omega}_{\mathcal{P}}$ as the associated $\mathcal{P}$-structures, an optimal strategy $U^{\star}$ must satisfy one of the following two assertions,*

- $U^{\star} \in \Pi_1 \cap \Pi'$;
- $U^{\star} \in \mathcal{C}(\hat{\Omega}_{\mathcal{P}}) \cap \Pi'$.

Proof. Theorem 1 implies $\Pi_1 \cup \mathcal{C}(\hat{\Omega}_{\mathcal{P}})$ contains all the solutions in Π that interdicts every critical structures. The above claim becomes evident when noticing that a feasible solution in this case must also belong to Π'. □

In the budget version, let Π be all the leader's strategies under an unlimited budget and Π' be the ones that satisfy the budget constraint. The above proposition entails that Algorithm 1 is still valid by precomputing an optimal solution from $\Pi_1 \cap \Pi'$ as an incumbent and restricting the problem to block all the $\mathcal{P}$-structures under the budget constraint. The remaining issue is that the authentic critical structure set $\hat{\Omega} = \{T \in \Omega \mid w(T) < z^{\star}\}$ is unknown since $z^{\star}$ is the optimal objective value of the budget version. The following proposition addresses this caveat.

Proposition 13. *Let $\hat{\Omega}^z := \{T \in \Omega \mid w(T) < z\}$, then $\hat{\Omega}^z_{\mathcal{P}} \subseteq \hat{\Omega}^{z'}_{\mathcal{P}}$ whenever $z \leq z'$.*

Proof. For $z \leq z'$, we have $\hat{\Omega}^z \subseteq \hat{\Omega}^{z'}$. Then, the statement is a direct consequence of the definition of the $\mathcal{P}$-structures and the order-reserving property of the cut operator $\mathcal{C}$. □

In particular, in each iteration of the cut generation procedure for the budget version problem, the incumbent objective value $\underline{z}$ from the sub-problem must be less than or equal to $z^{\star}$. Thus, the $\mathcal{P}$-structure generated with respect to $\underline{z}$ also belongs to $\hat{\Omega}_{\mathcal{P}} = \hat{\Omega}^{z^{\star}}_{\mathcal{P}}$. This shows that using Algorithm 1 with $\mathcal{P}$-structures generated with respect to $\hat{\Omega}^{\underline{z}}$ is still applicable for solving the budget version interdiction games.

7 Conclusion

This paper studies a particular class of supervalid inequalities to solve binary interdiction games. For an arbitrary bipartition of the leader's strategy space associated with a property $\mathcal{P}$, we identified a class of structures called the $\mathcal{P}$-structures, each of which induces a supervalid inequality for the corresponding problem. To design separation methods, we restricted our attention to the class of regular bipartition properties and derived a new characterization for the set of $\mathcal{P}_{\emptyset}$-structures, which results in an exact verification method. We further defined two special types of structures, the partial nuclei and the regenerable sets, based on which we developed a more efficient partial verifier. The classification of various types of structures also leads to a hierarchy map of the $\mathcal{P}_{\emptyset}$-structures for binary interdiction games.

We provided three general examples in which we apply our results to solve binary interdiction games targeting shortest paths, spanning trees, and vertex covers. Moreover, the realization of the

regenerable sets in these problems reveals interesting network structures, such as skeletons of a path decomposition and sociable vertices in a graph, which may deserve further investigation due to their unique properties.

Finally, we had shown that every feasible set of a greedoid is 0-regenerable in the corresponding greedoid interdiction game. Moreover, if this greedoid also satisfies the strong exchange property, the associated 0-core verification subroutine can be implemented as a greedy algorithm. Therefore, the separation procedure is guaranteed to be efficient in the greedoid interdiction games.

For future work, several promising directions exist to explore the application and extension of the proposed method. First, although the concept of $\mathcal{P}$-structures is defined for arbitrary bipartition properties $\mathcal{P}$, our current separation method primarily focuses on the null property. Investigating other types of bipartition properties to derive easily separable $\mathcal{P}$-structures, particularly by leveraging unique aspects of the follower's structures, presents an intriguing avenue. Second, our three examples demonstrate that the realization of k-regenerable sets varies significantly across problems, often revealing interesting network structures. This diversity invites further studies into the identification of regenerable sets in various contexts. Third, exploring different types of interactions between players and their potential to yield new kinds of supervalid inequalities is another promising area for investigation.

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Appendix

Computational Experiments on Supervalid Inequalities for Shortest Path Interdiction Derived from Path Decompositions

In this section, we conduct a computational study to provide some insights into the separation of supervalid inequalities derived from path decompositions and skeletons (see Section 4.3). These experiments also serve to analyze the effectiveness of such inequalities. We compare various computational performance aspects of the MCS formulation (2), as introduced in [59], against the results obtained by the MPS formulation (3) on a set of randomly generated instances. The computational experiments were conducted on a computer powered by a 12-core Intel Xeon E5-2620 v3 2.4 GHz processor, having 128 GB of RAM, and running Linux x86 64, CentOS 7.2. The formulations were implemented in Python and solved using the commercial optimizer Gurobi 11.0. Each instance was solved under a time limit of 30 minutes.

To conduct our experiments, we randomly generated a testbed of two-terminal series-parallel graphs (TTSPGs) [20] of different characteristics and sizes. A TTSPG is defined as a connected graph constructed by a sequence of series and parallel compositions starting from a set of copies of a single-edge graph K_2 (i.e., a clique of size two) with one terminal set as the source s and the other as the sink t. Given two TTSPGs G_1 and G_2, with sources and sinks s_1 and s_2, and t_1 and t_2, respectively, a series composition of G_1 and G_2 consists of a new graph G created by merging sink t_1 with source s_2, leaving s_1 and t_2 as the source and sink of G. In contrast, a parallel composition of G_1 and G_2 is a new graph G created by merging the sources s_1 and s_2 to produce the source of G and merging the sinks t_1 and t_2 to produce the sink of G.

The specific choice of graph type for these experiments is motivated by the fact that the vertex sets of all bi-connected components of a TTSPG directly induce a path decomposition, as described in Definition 10. This allows us to generate instances where multiple skeletons can be identified and tested to verify whether they induce one of the proposed supervalid inequalities. Also, the longest s-t path of a TTSPG can be identified in polynomial time, which allows us to verify the partial nuclei efficiently.

The graph generation was conducted as follows. First, we denote a *q-block* as a bi-connected TTSPG created by a sequence of parallel and series compositions of q single edge graphs K_2 so that the last composition is a parallel composition. This step ensures that the resulting q-block is bi-connected, thereby preventing the creation of trivial instances in which the attacker can disconnect the graph by interdicting just one edge. Specifically, we generate a random q-block by initializing a collection $\mathcal{Q}$ of q copies of K_2 and then progressively replacing two randomly selected graphs from $\mathcal{Q}$ with the result of either a series or a parallel composition chosen at random until only two graphs remain. Then, the resulting q-block is produced by applying a parallel composition to the remaining graphs in $\mathcal{Q}$.

We generated each instance for our experiments by sequentially applying series compositions to k randomly generated q-blocks. Consequently, each graph in our testbed is a randomly generated TTSPG composed of k bi-connected components (each being a q-block). To produce instances of different sizes, for each value of $k \in \{5, 10, 20\}$, we chose three different values of q so that the resulting sizes—measured in terms of the number of vertices and edges—can be roughly classified either as small, medium, or large. For each possible assignments of k and q, we generated five random instances with edge costs ($\{c_e\}_{e\in E}$) and edge lengths ($\{w_e\}_{e\in E}$) distributed uniformly in the interval $[1, 100]$.

In addition to the graph size, the complexity of the instances also depends on the values chosen for r, as this parameter dictates the number of s-t paths in $\hat{\Omega}$ that must be interdicted. We solve the problem for all generated graphs with five different levels for r. Specifically, we pick $r = (lp(G) - sp(G)) \times \gamma$, for $\gamma \in \{0.1, 0.2, 0.3, 0.4, 0.5\}$, where $lp(G)$ and $sp(G)$ stand for the lengths of the longest and shortest path and γ is referred to as the interdiction requirement. We call each triple (k, q, γ) an instance configuration. In total, the MCS and the MPS formulations were tested over 45 (graphs) $\times$ 5 (interdiction requirements) $= 225$ instances. Finally, since any minimum-cost s-t cut is a feasible solution for the attacker and is optimal in Π_1, we find one in the preprocessing step and feed it to the MCS formulation as an initial incumbent solution, as described in Algorithm 1.

Table 1 reports the results obtained for both the MCS and MPS formulations. For each instance configuration (k, q, γ), we show the average size of the five randomly generated instances in terms of the average number of vertices $\tilde{n}$ and edges $\tilde{m}$, the percentage of instances solved to optimality (Opt), the average run time in seconds (Time), the average optimality gap (Gap), and the number of cuts generated (Cuts). For the particular case of the MPS formulation, we specify the number of base cuts (i.e., the same type of cuts produced by the MCS formulation) and the number of supervalid cuts. We highlight in bold the best results obtained for each of these metrics.

From the data presented in the table, it is evident that MPS consistently outperforms MCS in most scenarios, particularly in larger instance sizes and when the interdiction requirement γ is high. A notable example is observed when (k, q) equals $(20, 40)$; here, MCS failed to solve any instances for $\gamma \geq 0.2$, whereas MPS successfully solved all. Furthermore, as γ increases, there is a significant reduction in the average computational times for MPS.

Even in cases where neither algorithms reached a solution, MPS generally achieved a smaller optimality gap. This is highlighted in the $(10, 150)$ instances where, at $\gamma \geq 0.2$, neither algorithm solved any instance. However, MPS displayed superior performance in terms of average optimality gaps, which also improved progressively with higher γ values.

An additional noteworthy point is the number of cuts generated by each algorithm. MPS demonstrated an enhanced capability to separate more supervalid constraints at higher γ values, leading to shorter solution times. This feature of MPS is further underscored in scenarios where MCS performed better; these instances coincided with situations where MPS produced no supervalid inequalities. Moreover, when MPS outperformed MCS, the total number of cuts it generated was

Configration			Avg. Size		Opt (%)		Time (s)		Gap (%)		Cuts MCS	Cuts MPS	
k	q	γ	$\tilde{n}$	$\tilde{m}$	MCS	MPS	MCS	MPS	MCS	MPS		base	sv
		0.1	73	114	**100**	**100**	**0.01**	**0.01**	**0**	**0**	17	17	0
		0.2	73	114	**100**	**100**	**0.03**	**0.03**	**0**	**0**	48	39	5
5	25	0.3	73	114	**100**	**100**	0.07	**0.03**	**0**	**0**	87	24	30
		0.4	73	114	**100**	**100**	0.09	**0.04**	**0**	**0**	114	9	42
		0.5	73	114	**100**	**100**	0.08	**0.03**	**0**	**0**	124	0	37
		0.1	438	667	**100**	**100**	0.58	**0.45**	**0**	**0**	251	251	0
		0.2	438	667	**100**	**100**	**41.91**	47.24	0	0	5,638	5,638	0
5	150	0.3	438	667	**80**	**80**	**626.93**	627.57	**20**	28	20,960	22,429	0
		0.4	438	667	**80**	**80**	713.74	**698.64**	**3**	4	32,460	30,003	324
		0.5	438	667	**100**	**100**	814.05	**270.58**	**0**	**0**	36,207	20,016	2,617
		0.1	587	892	**100**	**100**	**2.93**	2.97	**0**	**0**	1,308	1,378	0
		0.2	587	892	0	**20**	**903.86**	915.43	**44**	**44**	37,411	38,131	0
5	200	0.3	587	892	**20**	**20**	1,540.84	**1,495.4**	**27**	29	56,547	56,919	5
		0.4	587	892	**0**	**0**	**1,800**	**1,800**	27	**15**	60,283	53,676	2,631
		0.5	587	892	0	**40**	1,800	**1,254.06**	26	**8**	60,194	32,029	8,088
		0.1	283	447	**100**	**100**	0.17	**0.14**	**0**	**0**	97	97	0
		0.2	283	447	**100**	**100**	13.23	**9.53**	**0**	**0**	6,064	3,940	1,054
10	50	0.3	283	447	60	**80**	1,272.34	**827.86**	7	**3**	59,989	6,376	34,078
		0.4	283	447	60	**100**	1,283.67	**284.7**	8	**0**	57,913	1	27,952
		0.5	283	447	60	**100**	1,340.93	**81.72**	7	**0**	62,042	0	17,062
		0.1	592	896	**100**	**100**	**29.61**	29.95	**0**	**0**	5,471	5,312	0
		0.2	592	896	0	**20**	1,800	**1,589.33**	24	**23**	62,702	59,947	1,975
10	100	0.3	592	896	0	**20**	1,800	**1,420.78**	24	**17**	63,865	421	56,952
		0.4	592	896	0	**20**	1,800	**1,334.32**	25	**17**	61,618	0	52,013
		0.5	592	896	0	**20**	1,800	**1,252.16**	28	**8**	61,221	0	46,461
		0.1	890	1,351	**100**	**100**	**212.35**	288.75	**0**	**0**	22,482	22,961	0
		0.2	890	1,351	**0**	**0**	**1,800**	**1,800**	55	**48**	61,861	52,021	10,457
10	150	0.3	890	1,351	**0**	**0**	**1,800**	**1,800**	58	**36**	62,751	0	63,301
		0.4	890	1,351	**0**	**0**	**1,800**	**1,800**	55	**22**	62,662	0	56,385
		0.5	890	1,351	**0**	**0**	**1,800**	**1,800**	55	**21**	62,924	0	59,319
		0.1	293	456	**100**	**100**	5.53	**2.87**	**0**	**0**	3,058	760	1,483
		0.2	293	456	**100**	**100**	436.72	**48.65**	**0**	**0**	25,879	12	10,547
20	25	0.3	293	456	**100**	**100**	360.15	**10.21**	**0**	**0**	25,890	0	6,481
		0.4	293	456	**100**	**100**	215.89	**4.38**	**0**	**0**	25,889	0	3,891
		0.5	293	456	**100**	**100**	208.4	**2.51**	**0**	**0**	25,889	0	2,339
		0.1	458	720	0	**0**	**1,800**	**1,800**	8	**2**	77,145	29224	54,621
		0.2	458	720	0	**100**	1,800	**1,064.84**	8	**0**	75,699	0	58,562
20	40	0.3	458	720	0	**100**	1,800	**833.87**	12	**0**	61,267	0	38,925
		0.4	458	720	0	**100**	1,800	**201.48**	6	**0**	86,448	0	23,447
		0.5	458	720	0	**100**	1,800	**58.78**	8	**0**	77,122	0	8,935
		0.1	886	1,346	60	**100**	**1,064.08**	1,289.53	**0**	**0**	46,188	46,357	0
		0.2	886	1,346	**0**	**0**	**1,800**	**1,800**	**94**	103	62,182	59,119	161
20	60	0.3	886	1,346	**0**	**0**	**1,800**	**1,800**	97	**81**	61,363	0	61,018
		0.4	886	1,346	**0**	**0**	**1,800**	**1,800**	99	**83**	57,294	0	57,494
		0.5	886	1,346	**0**	**0**	**1,800**	**1,800**	99	**67**	55,835	0	53,974

Table 1: Computational results for the MCS and MPS formulations. For each graph configuration, we show the average size of the five randomly generated instances in terms of the average number of vertices $\tilde{n}$ and edges $\tilde{m}$, the percentage of instances solved to optimality (Opt), the average run time in seconds (Time), the average optimality gap (Gap), and the number of cuts generated (Cuts). For the particular case of the MPS formulation, we specify the number of base cuts (i.e., the same type of cuts produced by the MCS formulation) and the supervalid cuts.

significantly fewer than those by MCS. This underscores the effectiveness of supervalid inequalities in reducing the solution space and enhancing the lower bound of the problem.

In summary, this series of experiments demonstrates that incorporating the proposed supervalid inequalities into the MCS formulation yields a significant computational advantage. It is important to note that the proposed supervalid inequalities can also be incorporated as part of other solution approaches, like the dualize-and-combine method, as they can help strengthen the resulting formulations when solving shortest path interdiction on graphs with other topologies. For the particular

case of our computational experiments, given the unique nature of series-parallel graphs, the results highlight the potential benefits of our framework for solving some specific binary interdiction games.