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► **To cite this version:**

Julio Araujo, V. Campos, Frédéric Giroire, Leonardo Sampaio, R. Soares. On the hull number of some graph classes. [Research Report] RR-7567, 2011, pp.12. inria-00576581v1

HAL Id: inria-00576581

<https://inria.hal.science/inria-00576581v1>

Submitted on 14 Mar 2011 (v1), last revised 14 Sep 2011 (v2)

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N° 7567

Mars 2011

Domaine 2

 **R**
*apport
de recherche*

On the hull number of some graph classes. *

J. Araujo [†], V. Campos [‡], F. Giroire [†], L. Sampaio ^{† §}, R. Soares [†]

Thème : Algorithmique, programmation, logiciels et architectures
Équipe-Projet Mascotte

Rapport de recherche n° 7567 — Mars 2011 — 9 pages

Abstract: Given a graph $G = (V, E)$, the *closed interval* of a pair of vertices $u, v \in V$, denoted by $I[u, v]$, is the set of vertices that belongs to some shortest (u, v) -path. For a given $S \subseteq V$, let $I[S] = \bigcup_{u, v \in S} I[u, v]$. We say that $S \subseteq V$ is a *convex set* if $I[S] = S$.

The *convex hull* $I_h[S]$ of a subset $S \subseteq V$ is the smallest convex set that contains S . We say that S is a *hull set* if $I_h[S] = V$. The cardinality of a minimum hull set of G is the *hull number* of G , denoted by $hn(G)$.

We show that deciding if $hn(G) \leq k$ is a NP-complete problem, even if G is bipartite and we prove that $hn(G)$ can be computed in polynomial time for cactus and P_4 -sparse graphs.

Key-words: graph convexity, hull number, bipartite graph, cactus graph, P_4 -sparse graph.

* Research supported by the INRIA Equipe Associée EWIN.

[†] {julio.araujo, frederic.giroire, leonardo.sampaio_rocha, ronan.pardo_soares}@inria.fr - MASCOTTE Project, I3S (CNRS & UNS) and INRIA, INRIA Sophia-Antipolis, 2004 route des Lucioles, BP 93, 06902 Sophia-Antipolis Cedex France.

[‡] {campos@lia.ufc.br} ParGO Research Group - Universidade Federal do Ceará - UFC - Campus do Pici, Bloco 910. 60455-760 - Fortaleza, CE - Brazil.

[§] Partially supported by CAPES/Brazil and ANR Blanc AGAPE.

Le nombre enveloppe de quelques classes de graphes

Résumé : Étant donné un graphe $G = (V, E)$, l'intervalle fermé d'une paire de sommets $u, v \in V$, noté $I[u, v]$, est l'ensemble des sommets qui appartiennent à tous les plus courts chemins entre u et v . Pour un ensemble $S \subseteq V$, on note $I[S]$ l'ensemble $\bigcup_{u, v \in S} I[u, v]$. De plus, $S \subseteq V$ est dit *convexe* si $I[S] = S$.

L'*enveloppe convexe* $I_h[S]$ d'un sous-ensemble $S \subseteq V$ de G est défini comme le plus petit ensemble convexe qui contient S . $S \subseteq V$ est une *enveloppe* si $I_h[S] = V$. Le *nombre enveloppe* de G , noté $hn(G)$, est la cardinalité minimum d'une enveloppe de graphe G .

Nous montrons que décider si $hn(G) \leq k$ est un problème NP-complet, même si G est biparti et nous prouvons que $hn(G)$ peut être calculé en temps polynomial pour les cactus et pour les graphes P_4 -épars.

Mots-clés : convexité des graphes, nombre enveloppe, graphes bipartis, graphes P_4 -épars, graphes cactus

1 Introduction

In this paper, we adopt the graph terminology defined in [2]. All the graphs studied in this work are considered to be simple and undirected.

Given a graph $G = (V, E)$, the *closed interval* of a pair of vertices $u, v \in V$, denoted by $I[u, v]$, is the set of vertices that belong to some shortest (u, v) -path. If $S \subseteq V$, let $I[S] = \bigcup_{u, v \in S} I[u, v]$. We say that $S \subseteq V$ is a *convex set* if $I[S] = S$. Given a subset $S \subseteq V$, the *convex hull* $I_h[S]$ of S is the smallest convex set that contains S . We say that S is a *hull set* if $I_h[S] = V$. The cardinality of a minimum hull set of G is the *hull number* of G , denoted by $hn(G)$. Determining the hull number of an arbitrary graph is a NP-hard problem [7].

These definitions bring concepts from Euclidian space to Graph Theory. There are several works in the literature about these notions [3–9, 11].

In Section 3, we answer an open question of Dourado et al. [7] by showing that it is NP-hard to compute the hull number of a given bipartite graph. We then present polynomial time algorithms to compute the hull number of cacti, in Section 4, and for P_4 -sparse graphs in Section 5. This last result extends the previous known result for cographs [7]. Finally, we present some concluding remarks.

2 Preliminaries

In order to present our main results, it is necessary to first point some observations.

Lemma 1 ([7]). *Every simplicial vertex of a graph G belongs to any hull set of G .*

Lemma 2 ([7]). *If a graph G is not complete, then every universal vertex u of G does not belong to any minimum hull set of G .*

Lemma 3 ([7]). *Let G be a graph and H be an isometric subgraph of G . Then for every hull set S of H it holds that $V(H) \subseteq I_h(S)_G$.*

Lemma 4 ([7]). *Let G be a graph and S a proper and non-empty subset of $V(G)$. If $V(G) \setminus S$ is convex, then every hull set of G contains at least one vertex of S .*

Lemma 5. *Let $G = (V, E)$ be an arbitrary graph such that V can be partitioned into three sets K , a clique, R , a non-empty set which is not a clique, and $S = V \setminus (K \cup R)$. Moreover, for every vertex $v \in R$, $(N(v) \cap (V \setminus R)) = K$. Then, any minimum hull set of G contains a minimum hull set of $H = G[K \cup R]$.*

Proof. First, observe that either in G or in H if there is shortest (u, v) -path containing an internal vertex in R , then both u and v are, necessarily, in R because every vertex of R is neighbor of every vertex in K and of no vertex in S .

By the above remark, all vertices of R can only be generated by two vertices in R and every vertex of K is also generated by two non adjacent vertices of R . Consequently, to generate all the vertices of R in G , a subset $S_R \subseteq R$ must be chosen which is also a hull set of H . Therefore, any minimum hull set of G contains a minimum hull set of H . \square

3 Bipartite graphs

In this section, we study the complexity of deciding if $hn(G) \leq k$, for a bipartite graph G and an integer k given as input. However, we first show some approximability results.

Let $I(G)$ be the *incidence graph* of G , obtained from G by adding one vertex s_{uv} , for each edge $uv \in E(G)$, and replacing the edge uv by the edges $us_{uv}, s_{uv}v$.

Proposition 6. $hn(I(G)) \leq hn(G) \leq 2hn(I(G))$.

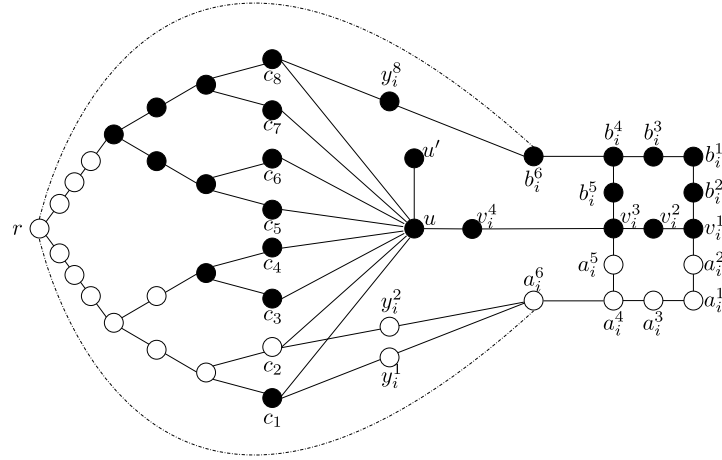


Figure 1: Subgraph of the bipartite instance $G(\mathcal{F})$ containing the gadget of a variable x_i that appears positively in clauses C_1 and C_2 and negatively in C_8 . White vertices belong to the set U in case C_2 is not satisfied.

Proof. Let $I(G)$ be the incidence graph of G . Observe that any hull set of G is a hull set of $I(G)$, since any shortest path in G is a shortest path in $I(G)$ (the edges were subdivided). Consequently, $hn(I(G)) \leq hn(G)$. However, given a hull set S_h of $I(G)$, one may find a hull set of G by simply replacing each vertex of S_h that represents an edge of G by its neighbors (vertices of G). Thus, $hn(G) \leq 2hn(I(G))$. \square

Corollary 7. *If there exists a k -approximation algorithm B to bipartite graphs, then B is a $2k$ -approximation algorithm for any graph.*

Due to the next result, one may not expect to find an exact polynomial time algorithm to calculate a hull number of a bipartite graph, unless $P = NP$.

Theorem 8. *Given a bipartite graph G and an integer k , deciding if $hn(G) \leq k$ is a NP-complete problem.*

Proof. This proof for bipartite graphs is just a slight modification of the proof for general graphs [7]. Further, it is no longer necessary to show that this problem is in NP, since it was already done for any graph.

We reduce the SATisfiability Problem to Hull Number in bipartite graphs. In this problem the input is a set of clauses $\mathcal{F} = \{C_1, C_2, \dots, C_m\}$ in the conjunctive normal form over a set of boolean variables $X = \{x_1, x_2, \dots, x_n\}$. The goal is to decide if there is an assignment of the variables such that each clause is satisfied.

We assume that m is a power of 2, since one can trivially add dummy variables and clauses without changing the satisfiability of \mathcal{F} . Moreover, we also assume that each variable x_i and its negation appear at least once in \mathcal{F} , otherwise the clauses that x_i appears could always be satisfied.

Given an instance \mathcal{F} to the SAT problem, we construct a bipartite graph $G(\mathcal{F})$ as follows. First, we build a full binary tree T rooted in r with m leaves, and we denote by $L = \{c_1, c_2, \dots, c_m\}$ the set of leaves of T . We then construct a graph H by adding a vertex u that is adjacent to every vertex in L and replace every edge joining a vertex $v \in V(T)$ to its parent in T by a path with $2^{h(v)}$ edges, where $h(v)$ is the distance of v to any of its decendent leaves.

The following claims are proved in [7].

Claim I. Let $v, w \in V(T) \setminus \{r\}$. The closed interval of v, w in H contains the parents of v in T if and only if v and w are siblings in T .

Claim II. The set L is a minimal hull set of H .

Finally, we build a graph $G(\mathcal{F})$ by adding to H the vertex u' whose only neighbor is u and by adding, for each variable x_i , a variable subgraph in $G(\mathcal{F})$ with vertex set $\bigcup_{j=1}^6 \{a_i^j, b_i^j\} \cup \{\bigcup_{j=1}^4 v_i^j\} \cup \{\bigcup_{\{j \mid x_i \in C_j\}} y_i^j\}$ and edges as depicted in Figure 1.

Furthermore, we complete the construction of $G(\mathcal{F})$ by adding, for every variable x_i , one edge joining u to v_i^4 and creating two paths with $2^{h(r)} - 3$ edges, one joining a_i^6 to r and the other joining b_i^6 to r . Denote by D the set of all internal vertices of these paths.

Claim III. $G(\mathcal{F})$ is a bipartite graph.

Let us present a proper 2-coloring c of $G(\mathcal{F})$. Let $c(r) = 1$, and for each vertex w in $V(H)$, define $c(w)$ as 1, if w is in an even distance from r , and 2 otherwise. Clearly, c is a partial *proper* coloring of $G(\mathcal{F})$ and moreover we have $c(u) = 1$ and $c(c_j) = 2$, for any $j \in \{1, \dots, m\}$. Let $c(u') = 2$. For each $i \in \{1, \dots, n\}$, let $c(y_i^j) = 1$, for every j such that $x_i \in C_j$, and $c(b_i^6) = c(a_i^6) = c(v_i^4) = 2$. Again, this partial coloring of $G(\mathcal{F})$ is proper. One can easily verify that this coloring can be extended to each variable subgraph. Moreover, since $c(r) = 1$ and $c(a_i^6) = 2$ ($c(b_i^6) = 2$), for every $i \in \{1, \dots, n\}$, and since the path that we add in $G(\mathcal{F})$ between r and a_i^6 (b_i^6) is of odd size, one can completely extend c in order to get a proper 2-coloring of $G(\mathcal{F})$.

Set $k = n + 1$. We show that \mathcal{F} is satisfiable if and only if the $hn(G(\mathcal{F})) \leq n + 1$.

First, consider that \mathcal{F} is satisfiable. Given an assignment A that turns \mathcal{F} true, define a set S as follows. For $1 \leq i \leq n$, if x_i is true in A add a_i^1 to S , otherwise add b_i^1 to S . Finally, add u' to S . Note that $|S| = n + 1$. We show that S is a hull set of $G(\mathcal{F})$. First note that $a_i^6, c_j \in I[a_i^1, u']$, for every clause C_j containing the positive literal of x_i . Similarly, observe that $b_i^6, c_j \in I[b_i^1, u']$, for every clause C_j containing the negative literal of x_i . Since A satisfies \mathcal{F} , it follows $L \subseteq I_h[S]$. Therefore, H being an isometric subgraph of $G(\mathcal{F})$, Lemma 3 and Claim II imply that $V(H) \subseteq I_h[S]$. Furthermore, the shortest paths between r and u have length $2^{h(r)}$, which implies that all vertices $a_{i,6}$ and $b_{i,6}$ are included in $I_h[S]$. It remains to observe that $I_h[a_i^6, b_i^6, w]$, where $w \in [a_i^1, b_i^1]$, contains the variable subgraph of x_i . Therefore we have that S is a hull set of $G(\mathcal{F})$.

To prove the sufficiency, we need the following claim:

Claim IV. The set $V(G(\mathcal{F})) \setminus \{a_i^1, a_i^2, v_i^1, b_i^1, b_i^2\}$ is convex, for any $i \in \{1, \dots, n\}$.

Denote $W = \{a_i^1, a_i^2, v_i^1, b_i^1, b_i^2\}$, for some i , and $W' = \{a_i^3, b_i^3, v_i^2\}$. It is easy to verify that for any pair $x, y \in W'$, $I[x, y]$ contains no vertex of W .

We prove the sufficiency by contradiction. Suppose that $G(\mathcal{F})$ contains a hull set S with at most $n + 1$ vertices and that \mathcal{F} is not satisfiable.

Clearly $u' \in S$, because u' is a simplicial vertex of $G(\mathcal{F})$. Furthermore, Claim IV and Lemma 4 imply that S must contain at least one vertex w_i of the set $\{a_i^1, a_i^2, v_i^1, b_i^1, b_i^2\}$, for every $i \in \{1, \dots, n\}$. Hence, $|S| = n + 1$. Since $v_i^1 \in I[u', a_i^1]$, $v_i^1 \in I[u', b_i^1]$, $a_i^2 \in I[u', a_i^1]$ and $b_i^2 \in I[u', b_i^1]$, we can assume, without loss of generality, that either $w_i = a_i^1$ or $w_i = b_i^1$, for every $i \in \{1, \dots, n\}$. Therefore S defines the following truth assignment A to \mathcal{F} . If $w_i = a_i^1$ set x_i true, else set x_i as false.

Using the hypothesis that \mathcal{F} is not satisfiable, we complete the proof by showing that there is a set U such that $V(G(\mathcal{F})) \setminus U$ is a convex set and $U \cap S = \emptyset$.

As \mathcal{F} is not satisfiable, there exists at least one clause C_j not satisfied by A . Let U' be the set of vertices belonging to the path from c_j to r in $G(\mathcal{F})$; and U'' formed by the vertices of $G(\mathcal{F})$ which are inner vertices of a path between a vertex from $U' \cap T$ and its child in T that is not in U' . Define $U \subset V(G(\mathcal{F}))$ as the union of the following sets (we suggest the reader to see Figure 1):

- the set D ;
- $U' \cup U''$;
- the union of the sets $\{a_i^p, 1 \leq p \leq 6\}$ (resp. $\{b_i^p, 1 \leq p \leq 6\}$) for each i such that x_i occurs positively (resp. negatively) in C_j ;

- the union of the sets $\{y_i^k \mid x_i \text{ appears positively in } C_k\}$ (resp. $\{y_i^k \mid x_i \text{ appears negatively in } C_k\}$) for each i such that x_i occurs positively (resp. negatively) in C_j .

Observe that $U \cap S = \emptyset$. It remains to prove that $V(G(\mathcal{F})) \setminus U$ is a convex set. In order to do this, we show that $I[w, w'] \cap U = \emptyset$ for every pair $w, w' \in V(G(\mathcal{F})) \setminus U$. Consider the following partition $A_1 \cup A_2 \cup A_3$ of $V(G(\mathcal{F})) \setminus U$:

- $A_1 = V(H) \setminus (U \cup \{u\})$;
- A_2 contains the vertices of the variable subgraphs which are not in U ;
- $A_3 = \{u, u'\}$

Case 1: $w, w' \in A_1$. We can assume that $w, w' \in V(T)$. Suppose by contradiction that there exists $x \in I[w, w'] \cap U$. It is easy to see that $x \notin D$, and, by using Claim I, $x \notin U \cap V(H)$. The only possibility is that $x = a_i^6$ or $x = b_i^6$, for some i . But in this case, there would be a strictly shorter path between w and w' passing through u .

Case 2: $w \in A_1$ and $w' \in A_2$. If w' does not belong to the gadget of a variable x_i that appears in C_j , then it is trivial to check that the shortest (w, w') -paths contain only vertices of $V(G(\mathcal{F})) \setminus U$. Otherwise, without loss of generality, we may assume that $w = c_k$, for some $k \in \{1, \dots, m\} \setminus \{j\}$ and that $w' \in \{v_i^1, v_i^3, v_i^4\} \cup \{y_i^l \mid \text{for every } l \text{ s.t. } y_i^l \notin U\}$. In any case, it is easy to check that $I[w, w']$ does not contain any vertex of U .

Moreover, all the other cases can be trivially verified. □

4 Cacti graphs

It is easy to observe that the hull number a tree is size of the set of its leaves. In this section, we generalize this simple result to the class of cacti. A *cactus* is a graph in which every pair of cycles have at most one common vertex. In order, to compute the hull number of a cactus, we present which vertices must be chosen.

Lemma 9. *Let C be an induced cycle in G . If C has at most one vertex v such that $d(v) > 2$, then any minimum hull set of G contains either one or two vertices of $V(C)$.*

Proof. The proof follows from the simple observation that $(V(G) \setminus V(C)) \cup \{v\}$ is a convex set. Thus, if C is odd, it is necessary to choose two vertices of C , otherwise, one suffices. □

Lemma 10. *Let C be an induced cycle in G . Let v_1, \dots, v_k be the vertices of C having degree greater than two, for $k \geq 2$. If there is a set of vertices S of C such that they are not in any minimum (v_i, v_j) -path, $1 \leq i, j \leq k$, in C , then any minimum hull set of G contains either one or two vertices of S .*

Proof. Again, observe that $V(G) \setminus S$ is a convex set and that S induces a path P_k on k vertices. If k is odd, then one vertex of S suffices to generate the others, otherwise, it is necessary to select two of them. □

Theorem 11. *Given a cactus G , $hn(G)$ can be found in linear time.*

Proof. By a simple breadth-first search all the cycles of a cactus can be detected. For each one, it is necessary to detect if it has at most one vertex of degree greater than two. In case it has at least two vertices of degree greater than two, it is necessary to also verify if there is a set S satisfying Lemma 10. Finally, every vertex of degree one must be chosen. These tests be done in linear time and the result follows by Lemmas 1, 9 and 10. □

5 P_4 -sparse graphs

A graph is said to be a *cograph* if it is P_4 -free. Computing the hull number of cographs can be done in polynomial time [7]. In this section, we extend this result to the class of P_4 -sparse graphs which properly contains the cographs.

A graph is P_4 -sparse if, for every $S \subseteq V$ such that $|S| = 5$, $G[S]$ contains at most one induced P_4 . Clearly, a cograph is a P_4 -sparse graph. We call G a *strict P_4 -sparse graph* if G is P_4 -sparse and contains at least one P_4 as induced subgraph.

The algorithm we present to compute the hull number of any P_4 -sparse graph G uses its modular decomposition tree $T(G)$. The idea is to perform a post-order traversal in $T(G)$ calculating $hn(G)$. The definition and classical results about the modular decomposition of graphs can be found in [12].

It is well-known that the neighborhood nodes of the modular decomposition of P_4 -sparse graphs are spiders [10].

A *spider* is a graph whose vertex set can be partitioned into disjoint sets S , K , and R such that: $|S| = |K| \geq 2$, S is a stable set, K is a clique; every vertex in R is adjacent to all the vertices in K and to no vertex in S ; there exists a bijection $f : S \rightarrow K$ such that either $N_G(s) \cap K = K - \{f(s)\}$, for all vertices $s \in S$, in this case the spider is called *thick spider*; or $N_G(s) \cap K = \{f(s)\}$, for all vertices $s \in S$, and it is called *thin spider*.

Note that R may be empty, it is the only possible non-trivial strong submodule of a spider, and that if $|S| = |K| = 2$, the spider is, by definition, thin and thick. However, in order to simplify the notation in the proofs, we consider that a spider is called *thick* if, and only if, $|S| = |K| \geq 3$. Moreover, we remark that it is possible to find a modular decomposition tree in such a way that the series and parallel nodes of the tree have only two children each.

Finally, we present how to compute the hull number of each inner node of the modular decomposition tree of a P_4 -sparse graph.

Lemma 12 ([7]). *If $G = G_1 \cup G_2$, for any two graphs G_1 and G_2 , then $hn(G) = hn(G_1) + hn(G_2)$.*

Lemma 13. *Let G_1 and G_2 be two P_4 -sparse graphs and $G = G_1 \oplus G_2$. Thus, $hn(G)$ can be computed in polynomial-time.*

Proof. If G_1 and G_2 are both cographs, then $hn(G)$ can be found in polynomial time according to Dourado et al. [7]. Let G_1 be a graph which is not complete and G_2 be a strict P_4 -sparse. We claim that two non adjacent vertices of G_1 form a minimum hull set of G because they generate all the vertices of G_2 and G_2 also have two non adjacent vertices that generate all vertices of G_1 .

It remains to consider the case in which G_1 is a complete graph and G_2 is a strict P_4 -sparse graph. First, observe that no vertex of G_1 belongs to any minimum hull set of G . Let r be the root of the modular decomposition tree of G_2 . Let us consider all the possible cases for r .

Case 1: r is a series node. We claim that in this case $hn(G) = hn(G_2)$.

As a matter of fact, since G_2 is composed by the join of graphs, the distance between any two vertices is at most 2. Consequently, any shortest (u, v) -path between vertices in G_2 is also a shortest (u, v) -path in G . Moreover, a minimum hull set from G_2 generates all vertices of $V(G_2)$ in G . Finally, since G_2 has two vertices u and v that are non-adjacent they generate every vertex of G_1 .

Case 2: r is a parallel node. Let $G_2 = \bigcup_{i=1}^k G_2^i$. In this case, we claim that $hn(G) = \sum_{i=1}^k f(i)$ such that

$$f(i) = \begin{cases} hn(G_1 \oplus G_2^i), & \text{if } G_2^i \text{ is not complete;} \\ |V(G_i)|, & \text{otherwise.} \end{cases}$$

We prove the claim by showing that any minimum hull set of G in this case is composed by all the vertices of G_2^i , for every i such that G_2^i is a complete graph, union all the minimum hull sets of $G_1 \oplus G_2^j$, for every j such that G_2^j is not complete.

Recall that, by Lemma 2, no vertex of G_1 belongs to a minimum hull set of G . By Lemma 1, every vertex of G_2^i such that G_2^i is complete must belong to every minimum hull set of G .

It remains to show that it is necessary and sufficient to also include a minimum hull set of every $hn(G_1 \oplus G_2^j)$ such that G_2^j is not complete, in every minimum hull set of G .

The necessity can be easily proved by just applying Lemma 5 to every G_2^i that is not a complete graph.

The sufficiency follows again from the fact that G_1 is generated by two non adjacent vertices of G_2 and for each G_2^i either $V(G_2^i)$ is chosen or $V(G_2^i)$ is generated by a minimum hull set of $G_1 \oplus G_2^i$.

Case 3: r is a neighborhood node. Thus, $G_2 = (S \cup K \cup R, E_2)$ is a spider. We claim that $hn(G) = hn(G_2) + 1$ if G_2 is a thin spider and $R = \emptyset$, otherwise $hn(G) = hn(G_2)$.

First, no matter if G_2 is a thin or a thick spider, by Lemma 1, every vertex of S must belong to any minimum hull set of G_2 and also of G and, again, no vertex of G_1 belong to any minimum hull set of G . Moreover, all vertices in S are simplicial in G_2 and by the same argument they must belong to any minimum hull set of G_2 .

Observe that, if G_2 is thin and $R = \emptyset$, S is a hull set of G_2 , but it is not a hull set of G since the vertices of K are not generated by them in G . Consequently, it is necessary and sufficient to add one, and just one, vertex of K to have a minimum hull set of G .

In the case that G_2 is thick and $R = \emptyset$, the vertices of S are sufficient to generate G_2 and also G , because every vertex of K is adjacent to at least two vertices of S that are necessarily non adjacent.

On the other hand, if $R \neq \emptyset$, then by Lemma 1 any minimum hull set of G_2 contains S and, by Lemma 5 a minimum hull set of $K \cup R$ which is composed by vertices of R .

By the same Lemmas, a minimum hull set of G_2 is a minimum hull set of G since, by Lemma 2, no vertex of G_1 belongs to any minimum hull set of G and G_1 is generated by non adjacent vertices of G_2 . \square

Lemma 14. *Let $G = (S \cup K \cup R, E)$ be a spider. Then,*

$$hn(G) = \begin{cases} |S| + hn(G[K] \oplus G[R]), & \text{if } R \neq \emptyset \text{ and } R \text{ is not a clique;} \\ |S| + |R|, & \text{otherwise.} \end{cases}$$

Proof. Since all the vertices in S are simplicial vertices, we apply Lemma 1 to conclude that they are all contained in any hull set of G .

By the structure of a spider, every vertex of K belongs to a shortest path between two vertices in S and are therefore generated by them in any minimum hull set of G . Consequently, if $R = \emptyset$, S is a minimum hull set of G . If R is a clique, $S \cup R$ is the set of simplicial vertices of G and also a minimum hull set of G .

Finally, if $R \neq \emptyset$ and R is not a clique, then Lemma 5 completes the proof. \square

Theorem 15. *Given a P_4 -sparse graph G , the hull number of G can be found in polynomial time.*

Proof. The modular decomposition tree of any graph G can be found in linear time [13].

The complexity of the post-order traversal is $O(n)$. For each inner node, the algorithm may spend $O(n^2)$ operations in case 2 of Lemma 13 or when $R \neq \emptyset$ in Lemma 14. Consequently, the complexity of the algorithm is $O(n^3)$. \square

6 Conclusions

In this paper, we simplified the reduction of Dourado et al. [7] to answer two questions they asked about the complexity of computing the hull number of triangle-free graphs and, in particular, of bipartite graphs. However, the complexity of this problem for interval graphs and chordal graphs remains open.

Finally, we used the modular decomposition tree of P_4 -sparse graphs to compute the hull number of these graphs. The P_4 -sparse graphs are contained in the class of $(q, q-4)$ -graphs for which every set of q vertices induce at most $q-4$ disjoint P_4 's [1]. It would be nice to extend our results to this class of graphs.

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ISSN 0249-6399