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Online graph coloring with bichromatic exchanges

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Abstract

Greedy algorithms for the graph coloring problem require a large number of colors, even for very simple classes of graphs. For example, any greedy algorithm coloring trees requires $\Omega(\log n)$ colors in the worst case. We consider a variation of greedy algorithms in which the algorithm is allowed to make modifications to previously colored vertices by performing local bichromatic exchanges. We show that such algorithms can be used to find an optimal coloring in the case of bipartite graphs, chordal graphs and outerplanar graphs. We also show that it can find colorings of general planar graphs with $O(\log \Delta)$ colors, where Δ is the maximum degree of the graph. The question of whether planar graphs can be colored by an online algorithm with bichromatic exchanges using only a constant number of colors is still open.

Keywords: online algorithms, graph coloring, bichromatic exchange, kempe chain.

1 Introduction

Online algorithms are a class of algorithms reading their input sequentially. In the case of graph problems, this usually means that the vertices of the graph arrive one by one. As the formal definition of greedy, online, and sequential algorithms is not completely fixed, we start by precisizing the convention we use here. In an online algorithm, for each new vertex, the algorithm must adapt a partial solution of the problem on the graph without the new vertex into a solution for the whole graph. For the graph coloring problem, this means that, if G is the graph and v the newly added vertex, the algorithm must transform a coloring of $G - v$ into a coloring of G .

A special class of online coloring algorithms which received a lot of attention is the class of greedy coloring algorithms, where the algorithm is not allowed to change his previous choices. The algorithm must assign a color to each new vertex that is different from the colors of its neighbors. Vertices colored at an earlier step cannot be recolored. The most common greedy algorithm for graph coloring is FIRST-FIT, where the color of the new vertex is taken to be the smallest color not already present in its neighborhood. These greedy algorithms are usually studied in a setting where the order on the vertices of the graph is arbitrary. The performance of these algorithms, is measured by the number of colors used for the worst case ordering.

A similar type of algorithms are sequential algorithms. These algorithms first decide on an ordering of the vertices, and then, apply an online algorithm to color the graph according to this ordering. In this case the ordering is chosen by the algorithm.

The greedy graph coloring problem is a widely studied subject. For general graphs, a randomized greedy algorithm finding an $O(\frac{n}{\log n})$ approximation was devised in [Hal97] while there is a lower bound of $\Omega(\frac{n}{\log^2 n})$ on the approximation ratio of any greedy algorithm. An important effort has been directed at studying the performance on usual graph classes. A greedy algorithm with an approximation ratio of $O(\log n)$ was shown for trees [GL88], bipartite graphs [LST89], planar and chordal graphs [Ira94]. On the other hand, this approximation ratio was shown to be optimal for these graph classes [Bea76, AS17, LST89]. The lower bounds even holds for randomized algorithms, algorithms using a small reordering

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buffer, or algorithm allowed to look at a few future inputs before making a choice [AS17]. Online algorithms with constant approximation ratio exists for other classes of graphs e.g. interval graphs [KT81, LV98, Smi10, NB08] and disk intersection graphs [CFKP07, EF02, AS17].

Since the approximation ratio of greedy algorithms is quite large even for some simple classes of graphs, it is natural to look at more general algorithms. In this article, we consider online algorithms which are allowed to change the color of previous vertices only by making local bichromatic exchanges. A bichromatic exchange (also called Kempe change in the litterature) consists in swapping the colors of the vertices of a maximal connected 2-colored subgraph of G . Applying this transformation creates a new proper coloring of the graph. We call these algorithms online algorithm with bichromatic exchanges. For each new vertex v , the algorithm can perform bichromatic exchanges to remove one color from the neighbors of v , and use this color for v .

The choice of bichromatic exchanges as an operation to recolor the graph is quite natural. It was first considered in Kempe’s failed attempt at proving the four color theorem, and was used later to prove Vizing’s theorem [Viz64]. More recently, sequential algorithm using bichromatic exchange were considered in [MP99] to color a special subclass of perfect graphs, and in [HG96, HGM98, HM97, Tuc87] using more complex recolorings. In this context, bichromatic exchanges are used to locally modify an existing coloring in order to color a new vertex.

Finally, bichromatic exchanges were considered in the context of graph coloring reconfiguration. In this case the problem is, given two colorings of the same graph, to decide whether one can be transformed into the other by applying a sequence of basic operations. The case where the allowed transformation is to recolor a single vertex (a particular case of bichromatic exchange) has attracted a lot of attention. Some results were found for specific graph classes, for example chordal graphs [BJL⁺14] and graphs with bounded treewidth [BB13, BB14]. Questions about the connexity of the reconfiguration graph [CVdHJ09, CVDHJ11] and some complexity aspects of the problem [BC09, CVDHJ11, JKK⁺14] have also been considered. The case where the recoloring operations are bichromatic exchanges has been considered for planar graphs [Mey78, Moh06], K_5 -minor free graphs [LVM81] and d -regular graphs [FJP15, BBFJ15].

This last problem corresponds to a more global setting where the bichromatic exchanges can be applied anywhere in the graph. In our case, we are interested in local modifications of the coloring. This means that, as in sequential algorithms, the bichromatic exchanges are restricted to the neighborhood of the newly added vertex.

Overview. The paper is organized as follows. In the next section, we give standard notations, and a formal definitions of online algorithm with bichromatic exchanges. Then, we describe such algorithms for several classes of graph: bipartite graphs in Section 2.1, chordal graphs in Section 3, outerplanar and planar graphs in Section 4. An extension of the algorithm on planar graphs to the case of graphs with bounded genus is presented in Section 5. Apart from the cases of planar graph and graphs with bounded genus, these algorithm are optimal and robust: they either find an optimal coloring or exhibit a forbidden substructure for this class of graphs. In the case of planar graph, the algorithm only provides a $\log(\Delta)$ coloring, where Δ is the maximum degree of the graph. The question of whether we can achieve a constant number of colors is still open. Finally, in Section 6 we give an example of a family of 3-colorable graph for which any online algorithm using bichromatic exchanges (with some additional restrictions) performs badly.

2 Definitions and notations

We start with some standard definitions and notations of usual concepts from graph theory. A graph is a pair (V, E) where V is the set of vertices, and E , the set of edges, is a subset of unordered pairs of distinct vertices. The graphs that we consider are simple, and undirected. Given a graph $G = (V, E)$ and a vertex v of G , we denote by $N_G(v)$ the neighborhood of v , i.e. $N_G(v) = \{u \in V, \{u, v\} \in E\}$, and $\Delta(G)$ the maximum degree of the graph G . When the graph G is clear from the context, we will omit the subscript and just write $N(v)$ the neighborhood of v , and Δ the maximum degree. If S is a subset of vertices, we denote by $G[S]$ the subgraph induced by all the vertices of S .

A k -coloring of a graph $G = (V, E)$ is a function $c : V \mapsto \{1, \dots, k\}$ such that there is no monochromatic edges: for every edge $\{u, v\} \in E$, $c(u) \neq c(v)$. Given a graph $G = (V, E)$, a coloring c of G , and a

color i , we denote by c_i the set of vertices colored i . In particular, for any two colors i and j , $G[c_i]$ is a stable set, and $G[c_i \cup c_j]$ is the bipartite graph induced by vertices colored i or j .

Given a graph G and two vertices u and v , a path between u and v is a sequence of distinct vertices $w_0 = u, w_1, \dots, w_k = v$ such that for all $0 \leq i < k$, w_i and w_{i+1} are adjacent in G . Given a coloring c of G , and two colors i and j , a path is a bichromatic (i, j) -path if every vertex of the path is colored either i or j .

We look at a particular recoloring procedure called bichromatic exchange. A bichromatic exchange is a transformation that changes a coloring into a new one. Informally this recoloring is done by selecting a vertex, changing its color to a new color, and propagating the change to its neighbors to prevent the creation of monochromatic edges. We give a formal definition below:

Definition 1. Let G be a graph, v a vertex of G , and j a color. We denote by $\langle v, j \rangle$ the bichromatic exchange that transforms a coloring c of G into a coloring $c' = \langle v, j \rangle(c)$ defined by:

$$c'(x) = \begin{cases} c(x) & \text{if } x \notin X \\ c(v) & \text{if } x \in X \text{ and } c(x) = j \\ j & \text{if } x \in X \text{ and } c(x) = c(v) \end{cases}$$

where X is the connected component of $G[c_{c(v)} \cup c_j]$ containing v .

It is easy to check that this transformation does not create any monochromatic edges. From the definition, we can see that the color of a vertex u is changed by the bichromatic exchange $\langle v, i \rangle$ if and only if there is a bichromatic $(i, c(v))$ -path between u and v . Let $S = \langle v_1, i_1 \rangle, \dots, \langle v_k, i_k \rangle$ be an ordered sequence of bichromatic exchanges, and c be a coloring. We denote by $\langle S \rangle(c) = \langle v_k, i_k \rangle \circ \dots \circ \langle v_1, i_1 \rangle(c)$ the coloring obtained after applying successively the bichromatic exchanges of S . In general, two bichromatic exchanges do not commute, i.e. the order of application of several bichromatic exchanges is important. However, here are a few simple cases where they do commute:

Remark 1. Let G be a graph, c be a coloring of G , and $\langle x, i \rangle, \langle y, j \rangle$ be two bichromatic exchanges.

- If the colors $i, j, c(x), c(y)$ are all different, then $\langle x, i \rangle \circ \langle y, j \rangle(c) = \langle y, j \rangle \circ \langle x, i \rangle(c)$.
- If $i = j$ and $c(x) = c(y)$, then $\langle x, i \rangle \circ \langle y, j \rangle(c) = \langle y, j \rangle \circ \langle x, i \rangle(c)$.

The algorithms that we consider are a special case of online algorithms where the algorithm is allowed to recolor previously colored vertices only by performing bichromatic exchanges. Allowing the algorithm to perform any bichromatic exchange would be too powerful as it would allow, in many cases, to recolor the whole graph. To prevent this, the algorithm is only allowed to perform bichromatic exchanges which are local to the newly added vertex in the following sense:

Definition 2. Let G be a graph, c a coloring of G , and v, x two vertices of G . The bichromatic exchange $\langle i, x \rangle$ is local to v if $x \in N(v)$.

We will always consider bichromatic exchanges which are local to v , the (not yet colored) vertex that was just added to the graph. Consequently, we will omit to specify v , and just write that the bichromatic exchanges are local as a shorthand for local to v . Note that vertices outside the neighborhood of v might be recolored by a local bichromatic exchange. Indeed, we only require that the starting point of a local bichromatic exchange is in the neighborhood of v , but this transformation can still recolor a large part of the graph. We can now give a formal definition of online algorithms with bichromatic exchanges.

Definition 3. An online coloring algorithm with bichromatic exchanges is an online algorithms such that, for each new vertex v :

- it applies a sequence of local bichromatic exchanges,
- and then selects a color for v not used by any vertex in its neighborhood.

This definition is quite general, however the algorithms that we will consider are more simple: they only perform bichromatic exchanges when necessary. If a color is already available to the new vertex without any recoloring, then the algorithm immediately selects that colors. Additionally, they always choose the smallest color available, and make no assumption on eventual properties of the existing

coloring. In other words, with these restrictions an algorithm with bichromatic exchanges is given by a procedure taking as input a graph G with a vertex v and any coloring c of $G - v$, and finding a sequence of bichromatic exchanges such that applying these transformations removes one color from the neighborhood of v .

In terms of graph coloring reconfiguration, the question this type of algorithms try to answer can be formulated in the following way. Given a graph G , a vertex v , and a k -coloring c of $G - v$, is there a transformation of c using local bichromatic exchanges into a coloring c' for which $N(v)$ is $(k - 1)$ -colored. Thus, the problem is similar to a coloring reconfiguration problem with two main variations:

- the bichromatic exchanges must be local,
- the target coloring is not given, but can be any coloring with the desired properties.

An other important parameter for this kind of algorithm is the length of the reconfiguration path: the number of bichromatic exchanges that must be applied for each new vertex in the worst case. Since we are interested in only making local changes to the coloring, we want this number to be polynomial in Δ .

The question we are interested in is to determine for which classes of graphs does such an algorithm exists and how many colors it requires. In the following, we will show several algorithms for different classes of graphs. Moreover, the algorithms that we will describe are robust in the following sense. For any graph G , if we run the algorithm A on G , then:

- either $G \in \mathcal{C}$, and A finds a coloring of G ,
- or, A exhibits a proof that $G \notin \mathcal{C}$.

2.1 Bipartite graphs

The first simple class of graph that we might look at are bipartite graphs. We prove the following:

Theorem 2. *There is an online coloring algorithm with bichromatic exchange finding a 2-coloring of bipartite graphs using at most Δ operations at each step.*

Proof. Let G be the graph, v a vertex of G , and c the coloring of $G - v$ obtained from the previous steps of the algorithm. If one of the two colors is not present in the neighborhood of v , then this color can be used to color v . Otherwise, the recoloring procedure is the following: while there is a vertex $u \in N(v)$ colored 2, apply the bichromatic exchange $\langle u, 1 \rangle$. None of the vertices in $N(v)$ colored 1 can change color back to 2. Indeed, assume by contradiction that at some step a vertex $w \in N(v)$ changes color from 1 to 2 during the bichromatic exchange $\langle u, 1 \rangle$. This implies that there is a path p of odd length from u to w in $G - v$. Consequently, $p \cup \{v\}$ is an odd cycle, a contradiction of the assumption that G is bipartite.

When the procedure ends, after at most Δ bichromatic exchanges, all the vertices in $N(v)$ are colored 1, and v can be colored 2. \square

In particular, this algorithm colors optimally any tree, while the best greedy online algorithm needs $\Omega(\log n)$ colors in the worst case.

3 Chordal graphs

In this section, we will exhibit an online algorithm with bichromatic exchanges that can color optimally any chordal graph. A chordal graph is a graph with no induced cycles of length larger than 3. They have the following property:

Theorem 3 (Dirac, 1961). *A graph G is chordal if and only if there is an ordering u_1, \dots, u_n of its vertices such that for all $i \geq 1$, $N(u_i) \cap \{u_1, \dots, u_{i-1}\}$ is a clique. Such an ordering is called an elimination ordering.*

Such ordering can be computed in polynomial time. Chordal graphs are an important subclass of perfect graphs for which the chromatic number is equal to the largest clique in the graph. The idea of the algorithm is the following. Let v be the new vertex that was added to the graph. The algorithm will recolor $N(v)$ using one less color. Since the subgraph induced by $N(v)$ is also chordal, we can find

an elimination ordering of the vertices in $N(v)$, and recolor these vertices one by one, according to this ordering. The property that the whole graph is chordal will ensure that the colors of previously recolored vertices do not change during the successive operations. We first prove the following lemma:

Lemma 4. *Let G be a graph, v a vertex of G , and c a coloring of $G - v$. Assume that there are two vertices $u_1, u_2 \in N(v)$, such that there is a bichromatic (i, j) -path between u_1 and u_2 in $G - v$. If p is the shortest such path, then $p \subseteq N(v)$.*

Proof. Let u_1 and u_2 be two neighbors of v such that there is a bichromatic (i, j) -path between u_1 and u_2 in $G - v$. Take p to be the shortest bichromatic (i, j) -path from u_1 to u_2 for some colors i and j . We will show that $p \subseteq N(v)$. Assume by contradiction that this is not the case, and there is at least one vertex on the path p which is not a neighbor of v . We can find a subpath p' of p such that p' has length at least 2, and the only vertices of p' in $N(v)$ are its two endpoints.

Since p is the shortest bichromatic (i, j) -path from u_1 to u_2 , the graph induced by p' is a path, and consequently, $G[v \cup p']$ is a cycle of length at least 4. This is a contradiction of the assumption that G is chordal. \square

Theorem 5. *There is an online coloring algorithm with bichromatic exchange that colors optimally every chordal graph. The algorithm performs at most Δ bichromatic exchanges for each new vertex.*

Proof. Given a graph H , we denote by $\omega(H)$ the size of the largest clique in H . Let G be a graph, v a vertex of G , and c a coloring of $G - v$ using $\omega(G - v)$ colors. If $\omega(G) = \omega(G - v) + 1$, then we can directly color v using the new additional color. Otherwise, let $\omega = \omega(G)$. We will describe an algorithm finding a sequence of bichromatic exchanges such that the resulting coloring uses at most $\omega - 1$ colors on $N(v)$. The transformation is made using at most Δ bichromatic exchanges.

Since G is chordal, the induced subgraph $G[N(v)]$ is also chordal. Let $k = |N(v)|$ be the number of neighbors of v , and let v_1, \dots, v_k be an elimination ordering of the vertices of $N(v)$. We write G_i the subgraph induced by the vertices v_1, \dots, v_i . The recoloring procedure is, for each i from 1 to k , if v_i is colored ω , we apply the bichromatic exchange $\langle v_i, x_i \rangle$ where x_i is the smallest color not present in $N_{G_i}(v_i)$. There is always such a color x_i . Indeed, by definition of the order of the vertices, $N_{G_i}(v_i)$ is a clique, and $N_{G_i}(v_i) \cup \{v_i, v\}$ is also a clique of G . Consequently, $|N_{G_i}(v_i)| \leq \omega - 2$.

During step i , none of the vertices v_j with $j < i$ is recolored. Indeed, suppose by contradiction that this is not the case, and let i be the first step at which a vertex v_j , with $j < i$ is recolored. Before the exchange, we have $c(v_j) \neq \omega$. After the exchange, the color of v_j changes to ω . Consequently, there is bichromatic (ω, x_i) -path from v_i to v_j in $G - v$. By Lemma 4, this implies that there is a bichromatic (ω, x_i) -path $p \subseteq N(v)$ from v_i to v_j .

Let v_k be the vertex of the path p with the largest index k . By choice of the color x_i , we know that $v_k \neq v_i$, since the (only) neighbor of v_i in p has an index larger than i . Consequently, since $j < i$, the vertex v_k has two neighbors v_a and v_b in p . Since p is 2-colored, both v_a and v_b have the same color. Additionally, by choice of k , both a and b are smaller than k . Since the ordering of the vertices is an elimination ordering, the neighbors of v_k in G_{k+1} form a clique. In particular, there is an edge between v_a and v_b , and this edge is monochromatic, a contradiction.

After the recoloring is done, none of the vertices of $N(v)$ is colored ω , and the color ω can be assigned to v . \square

4 Planar graphs

The goal of this section is to show two online algorithm with bichromatic exchanges. The first one allows to color outer-planar graphs using at most 3 colors. The second one can color any planar graph, but only computes a $O(\log \Delta)$ coloring. We start by the algorithm on outer-planar graphs.

4.1 Outer-planar graphs

An outerplanar graph is a graph which admits a planar drawing such that there is a face adjacent to all the vertices of the graph. Such a drawing of an outerplanar graph can be computed in linear time [Bre77]. The face containing all the vertices of the graph is the outer-face. Given an outerplanar graph G , and an outer-planar drawing of G , we can consider the sequence of vertices in the order they appear on the

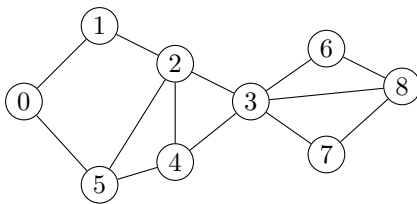


Figure 1: Example of outer-planar graph. The vertex 3 appears twice on the outer face of the graph.

outer face. Note that the same vertex can appear several times in this sequence as in the example in Figure 1. We will prove the following Theorem:

Theorem 6. *There is an online coloring algorithm with bichromatic exchange that colors optimally any outer-planar graph. The algorithm performs at most Δ bichromatic exchanges at each round.*

Proof. Let G be an outer-planar graph, v a vertex of G , and c an optimal coloring of $G - v$. If G is bipartite, then we can use the procedure from Theorem 2 to remove one color from $N(v)$, and obtain a 2-coloring of G . Otherwise, we can assume that the colors 1, 2 and 3 are present in $N(v)$.

We will describe a procedure recoloring $N(v)$ with only two colors using at most Δ local bichromatic exchanges.

Let $k = |N(v)|$, and let $v_0 = v, v_1, \dots, v_k$ be an ordering of the vertices of $N(v) \cup v$, in the order they appear on the outer face in an outerplanar drawing of G . Note that there is no ambiguity in the choice of this ordering. Indeed, suppose that there is a vertex, say v_i , that appears several times on the outer-face. Then none of the vertices that appear between the first and last occurrence of v_i on the outer face can be neighbors of v , since it would contradict the fact that G is outer-planar. We will recolor the vertices v_1, \dots, v_k successively using the colors 1 and 2.

The algorithm recolors the vertices v_1, \dots, v_k in this order by doing the following. For i from 1 to k , if $c(v_i) = 3$, let $x \in \{1, 2\} \setminus \{c(v_{i-1})\}$ (if $i = 1$, x can be either 1 or 2). We apply the bichromatic exchange $\langle v_i, x \rangle$.

During step i , none of the vertices v_1, \dots, v_{i-1} are recolored. Indeed, since G is outerplanar, any path in $G - v$ from v_i to some vertex v_j with $j < i$ necessarily goes through v_{i-1} . Additionally, the color of v_{i-1} is different from x and 3 by choice of x . Consequently, v_{i-1} is not contained in any bichromatic $(x, 3)$ -path, and there is no bichromatic $(x, 3)$ -path between v_i and v_j . At the end of this procedure, $N(v)$ is colored with colors 1 and 2, and v can be colored 3. \square

4.2 General planar graphs

The case of general planar graph is more complicated, and we will only prove the following weaker theorem.

Theorem 7. *There is an online algorithm with bichromatic exchanges coloring any planar graph with $O(\log(\Delta))$ colors. The algorithm performs at most $\frac{\Delta^2}{2}$ bichromatic exchanges for each new vertex.*

The question of whether there exists an algorithm using only a constant number of colors is still open. The algorithm is quite simple but the analysis is more complex than previous cases. The idea is to repeatedly apply a greedy procedure trying to remove one color from the neighborhood of the new vertex v . We will show that when this greedy procedure cannot be applied any more, $N(v)$ is colored using at most $O(\log \Delta)$ colors.

We start by a few definitions. Let G be a graph, v a vertex of G , and c a coloring of $G - v$. Given a color i , the *size* of i in the coloring c is the number of neighbors of v colored i : $\text{size}_c(i) = |N(v) \cap c_i|$. Let $\langle u, j \rangle$ be a local bichromatic exchange, and let $c' = \langle u, j \rangle(c)$ and $i = c(u)$. Suppose that $\text{size}_c(i) \leq \text{size}_c(j)$. The bichromatic exchange $\langle u, j \rangle$ is said to *increase inequalities* in c if we have $\text{size}_{c'}(i) < \text{size}_c(i)$ and $\text{size}_{c'}(j) > \text{size}_c(j)$. Intuitively, a bichromatic exchange increases inequalities if it decreases the size of colors with smaller sizes, and increases the size of color with larger sizes. The greedy recoloring procedure that we apply is the following.

Procedure (GreedyRecolor). *While there is a local bichromatic exchange $\langle u, i \rangle$ that increases inequalities, apply this bichromatic exchange.*

To prove the theorem, we only need to prove two things: (i) the procedure **GreedyRecolor** ends in a polynomial in Δ number of rounds, and (ii) if c is a coloring such that there is no local bichromatic exchange increasing inequalities, then $N(v)$ is colored with at most $O(\log \Delta)$ colors. The first point is proved in the following Lemma 8. The second point will be proved in Lemma 10 in the next subsection.

Lemma 8. *The procedure **GreedyRecolor** ends after at most $\frac{\Delta^2}{2}$ rounds.*

Proof. To show this result, we will exhibit a potential function Ψ such that:

- Ψ increases at every iteration of **GreedyRecolor**,
- Ψ is upper bounded.

Let c be a K -coloring, we define the potential $\Psi(c)$ by:

$$\Psi(c) = \sum_{\text{color } i} (\text{size}_c(i))^2.$$

We will show that Ψ increases by at least 2 at each iteration of **GreedyRecolor**. Let $\langle u, j \rangle$ be a local bichromatic exchange increasing inequalities in c , and let $i = c(u)$ and $c' = \langle u, j \rangle(c)$. By assumption on $\langle u, j \rangle$, we know that $\text{size}_{c'}(i) \leq \text{size}_c(i)$. Moreover, there is an integer x such that:

- $\text{size}_{c'}(i) = \text{size}_c(i) - x$
- $\text{size}_{c'}(j) = \text{size}_c(j) + x$

Since $\langle u, j \rangle$ increases inequalities, we have $x \geq 1$. Additionally, the following holds:

$$\begin{aligned} \Psi(c') - \Psi(c) &= (\text{size}_{c'}(i))^2 + (\text{size}_{c'}(j))^2 - (\text{size}_c(i))^2 - (\text{size}_c(j))^2 \\ &= (\text{size}_c(i) - x)^2 + (\text{size}_c(j) + x)^2 - (\text{size}_c(i))^2 - (\text{size}_c(j))^2 \\ &= 2x^2 + 2x(\text{size}_c(j) - \text{size}_c(i)) \geq 2 \end{aligned}$$

Moreover $\Psi(c)$ is upper bounded by Δ^2 for any coloring c . Indeed, since we know that for any coloring c , $\sum_i \text{size}_c(i) = \Delta$, $\Psi(c)$ is maximum when $\text{size}_c(i)$ is zero for all but one color. The potential Ψ is positive, upper bounded by Δ^2 , and increases by 2 at each iteration of **GreedyRecolor**. Consequently, the procedure **GreedyRecolor** must end after at most $\frac{\Delta^2}{2}$ rounds. \square

4.3 Grundy coloring of circular graphs

To prove the correctness of the algorithm described in previous subsection, we only need to show that if a coloring has no local bichromatic exchange that increases inequalities, then $N(v)$ is colored using at most $O(\log \Delta)$ colors. To prove this, we will use a result on a particular drawing of a graph which could be of independent interest. We will first describe this construction, and prove that any coloring satisfying some properties related to this drawing uses a small number of colors. This will allow us to complete the proof of Theorem 7.

Let G be a graph, with G not necessarily planar. A drawing of G is a function f that associates to each vertex v of the graph a point $f(v)$ in the plane, and to each edge (u, v) of the graph a path from $f(u)$ to $f(v)$. We call a *circular drawing* of G a drawing of G such that all the vertices are represented by points on a circle, and the paths representing the edges are straight line segments.

Definition 4. *Let $G = (V, E)$ be a graph, with a drawing of G . A proper coloring c of G is intersection compatible if for any two crossing edges (u_1, v_1) and (u_2, v_2) , the two sets $\{c(u_1), c(v_1)\}$ and $\{c(u_2), c(v_2)\}$ intersect.*

An example of a circular drawing of a graph with an intersection compatible coloring is given in Figure 2.

Not all graphs G have a drawing with an intersection compatible coloring. Determining which graphs have one could be an interesting question on its own. For example, there is no intersection compatible coloring of K_5 , the clique on 5 vertices, for any of its drawing on a plane. On the contrary, any 3-colorable

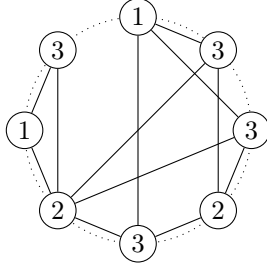


Figure 2: Example of circular drawing of a graph with an intersection compatible coloring.

graph has one. In fact any 3-coloring of a 3-colorable graph is intersection compatible for any drawing of the graph, since in this case any two edges share at least one color.

Given a graph G with a coloring c , the colored graph (G, c) will be called circular if there is a circular drawing of G that makes the coloring c intersection compatible. A k -coloring c of a graph is a *Grundy k -coloring* if for every vertex u , and for every color $j < c(u)$, there is a neighbor $w \in N(u)$ such that $c(w) = j$, and at least one vertex is colored k . In other words, every vertex is assigned the smallest color that is not present in its neighborhood. The Grundy chromatic number of a graph G is the largest number of colors k such that G has a Grundy k -coloring. To put it differently, a coloring is a Grundy coloring if it can be obtained from the algorithm FIRST-FIT, with a certain ordering of the vertices. The Grundy chromatic number is the largest number of colors that the algorithm FIRST-FIT might need for the worst case ordering. We will prove the following result.

Lemma 9. *There is a constant K_0 , such that for any circular colored graph (G, c) with n vertices, if c is a Grundy k -coloring of G , then $k \leq K_0 \log n$.*

Note that the lemma does not prove that the Grundy chromatic number of a circular graph is at most $O(\log n)$. Indeed, a circular graph could have a Grundy coloring using more colors. However, the lemma states that in this case, this coloring is not intersection compatible.

Proof. We denote by $T(k)$ the smallest number of vertices of a circular graph G having an intersection compatible Grundy k -coloring c . To prove the lemma, it is enough to show that there is a constant $\alpha > 1$ such that $T(k) \geq \alpha^{k-1}$ for any $k \geq 1$. We take $\alpha = 2^{\frac{1}{4}}$, and show this inequality by induction on k . If $k \leq 4$, we know that $T(k) \geq k \geq \alpha^{k-1}$.

Let us assume that $k > 4$, and let (G, c) be a colored circular graph with $T(k)$ vertices such that c is a Grundy k -coloring. Let u_0 be a vertex colored k . Since c is a Grundy coloring, there are three vertices u_1, u_2 , and u_3 adjacent to u_0 such that for each $i \in \{1, 2, 3\}$, u_i is colored $k - i$. Let $v_0 = u_0, v_1, v_2, v_3$ be the vertices $\{u_0, u_1, u_2, u_3\}$ in the order they appear on the circle in the corresponding circular drawing of G . We denote by G' the subgraph induced by the vertices with a color at most $k - 4$. We will show that G' is composed of at least 2 connected components, each of which contains all the colors from 1 to $k - 4$. By applying the induction hypothesis on each of these components, this gives as needed:

$$T(k) = |G| \geq |G'| \geq 2 \cdot T(k - 4) \geq 2 \cdot \alpha^{k-5} = \alpha^{k-1}.$$

Thus, we only need to show that G' contains at least two connected components, each of which contains all the colors from 1 to $k - 4$. Since c is a Grundy coloring and $c(v_1) > k - 4$, there is a vertex w_1 adjacent to v_1 with $c(w_1) = k - 4$. Similarly, there is a vertex w_2 adjacent to v_3 with color $c(w_2) = k - 4$. See Figure 3 for an illustration. Let G_1 and G_2 be the connected components of G' containing w_1 and w_2 respectively. Then $G_1 \neq G_2$. Indeed, if w_1 and w_2 were in the same connected component of G' , then there would be a path from w_1 to w_2 using only colors less than or equal to $k - 4$. By adding the two edges (v_1, w_1) and (v_3, w_2) , we would obtain a path from u_1 to u_3 which does not use the colors $k = c(v_0)$ and $c(v_2)$. However, this path must necessarily cross the edge (v_0, v_2) , a contradiction of the assumption that the coloring c is intersection compatible.

Consequently we have $T(k) \geq \alpha^{k-1}$ which can be rewritten as $k \leq 1 + \frac{\log(T(k))}{\log(\alpha)} \leq K_0 \log(n)$ with $K_0 = 1 + \frac{1}{\log(\alpha)}$. \square

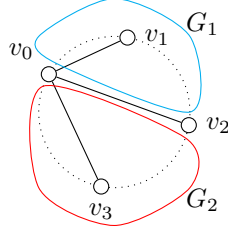


Figure 3: Any edge crossing the edge (v_0, v_2) must have either $c(v_0)$ or $c(v_2)$ as the color of one of its endpoints. Since G' contains no vertex with color $c(v_0)$ or $c(v_2)$, G' is split into two connected components, one on each side of the edge (v_0, v_2) .

We can now see how to use the previous result to complete the proof of the algorithm on planar graphs. The following lemma is the only remaining missing part to prove Theorem 7.

Lemma 10. *Let G be a planar graph, v a vertex of G , and c a coloring of $G - v$. If there is no local bichromatic exchange that increases inequalities, $N(v)$ is colored using at most $K_0 \log \Delta$ colors for some constant K_0 .*

Proof. We build the graph G' as follows. The set of vertices of G' is $N(v)$, the neighbors of v . For every pair of vertices $u_1, u_2 \in N(v)$ such that $c(u_1) \neq c(u_2)$, we add the edge (u_1, u_2) in G' if there is a bichromatic path from u_1 to u_2 in G . We consider a circular drawing of G' where the vertices appear on the circle in the order they appear around the vertex v in a planar drawing of G .

The coloring c of G induces a coloring c' of G' such that:

- The coloring c' is proper. Indeed, we only added edges in G' between vertices with different colors.
- The coloring c' is intersection compatible. Let (u_1, u_3) and (u_2, u_4) be two crossing edges in G' . Then, the vertices u_i appear with the order u_1, u_2, u_3, u_4 around the vertex v in the planar drawing of G . Additionally, there is a bichromatic $(c(u_1), c(u_3))$ -path p_1 in G from u_1 to u_3 , and a bichromatic $(c(u_2), c(u_4))$ -path p_2 from u_2 to u_4 . Since G is planar, p_1 and p_2 must cross at some vertex w . This implies that $c(w) \in \{c(u_1), c(u_3)\} \cap \{c(u_2), c(u_4)\}$, and consequently, this intersection is not empty.
- Finally, assume that the colors are ordered $1, \dots, k$ such that $\text{size}_c(1) \geq \text{size}_c(2) \geq \dots \geq \text{size}_c(k)$. Then c is a Grundy coloring of G' for this ordering of the colors. Indeed let u be a vertex of G' , and assume by contradiction that there is a vertex u and a color $i < c'(u)$ such that u has no neighbor colored i . Then consider the bichromatic exchange $\langle u, i \rangle$ in the graph G . This bichromatic exchange does not change the color of any neighbor of v colored i . And since we have $i < c(u)$, this implies $\text{size}(i) \geq \text{size}(u)$. Performing the bichromatic exchange $\langle u, i \rangle$ would increase the size of i , and decrease the size of $c(u)$. Thus $\langle u, i \rangle$ increases inequalities, a contradiction of the assumption that c did not contain any bichromatic exchange increasing inequalities.

By applying Lemma 9 on the graph G' with the coloring c' , we obtain that there is a constant K_0 such that c' uses at most $K_0 \log \Delta$ different colors. Consequently, the coloring c uses at most $K_0 \log \Delta$ colors on $N(v)$. \square

We now have all the ingredients to prove the theorem.

Proof of Theorem 7. Let K_0 be the constant in Lemma 10. The algorithm uses $K_0 \log \Delta + 1$ colors. For each new vertex v , the procedure **GreedyRecolor** is applied on $N(v)$. By Lemma 8, the procedure performs at most $\frac{\Delta^2}{2}$ of bichromatic exchanges. After the procedure has ended, we obtain a coloring c' such that there is no local bichromatic exchange that increases inequalities. By Lemma 10, this implies that c' colors $N(v)$ using at most $K_0 \log \Delta$ colors. Consequently, the algorithm colors any planar graph with $K_0 \log \Delta + 1$ colors. \square

5 Graphs of bounded genus

In this section, we show that the algorithm we described above can be adapted to the case of graphs with bounded genus, by only paying a small additional number of colors. We start by some very quick notations. We will assume that the reader is familiar with basic notions of topology and surfaces with boundaries. An introduction to these notions can be found for example in chapter 4 from [Ada04].

Given a triangulation T of a surface S , we denote by $V(T)$, $E(T)$ and $F(T)$ the set of vertices, edges and faces respectively of the triangulation T . Let S be a surface with a triangulation T . The *Euler characteristic* of S is the quantity $|V(T)| - |E(T)| + |F(T)|$. This quantity is an invariant of the surface S that is preserved by homeomorphism, and is independent of the chosen triangulation. The Euler characteristic can be negative, and is at most 2 for a connected closed surface, and at most 1 for a connected surface with boundary. For an orientable surface S , the genus g of S is related to its Euler characteristic h by the formula $h = 2 - 2g$.

A drawing of a graph G on a surface S is a mapping f that associates to every vertex v of G , a point $f(v)$ on S , and to every edge (u, v) of G , a path on S from $f(u)$ to $f(v)$. An embedding of a graph G on S is a drawing f of G on the surface S such that for any two edges e_1 and e_2 , the paths $f(e_1)$ and $f(e_2)$ do not intersect. In the following, the Euler characteristic (resp. genus) of a graph G will denote the largest number h (resp. smallest number g) such that there exists an embedding of G on a connected surface S with Euler characteristic h (resp. genus g). Planar graphs have Euler characteristic 2 and genus 0. The goal of this section is to show the following theorem:

Theorem 11. *There is an online algorithm with bichromatic exchanges that can color any graph with Euler characteristic h using $5 - 2h + O(\log \Delta)$ colors.*

The algorithm is exactly the same as for the planar case. To remove one color from the neighborhood of a vertex v , we just apply the procedure **GreedyRecolor**. Thus, to prove the correctness of the algorithm, we only need to show that, when there is no more bichromatic exchanges increasing inequalities, the neighborhood of the v is colored using at most $4 - 2h + O(\log \Delta)$ colors.

To prove this result, we generalize the definition of circular graph from previous section to handle graphs drawn on surfaces with higher genus. Let S be a surface with boundary, and G be a graph. A boundary drawing of G on S is a drawing f of G on the surface S such that for every vertex v , $f(v)$ is on the boundary of S . In a boundary drawing, paths corresponding to different edges are allowed to intersect. In particular, if S is a disk, a boundary drawing of a graph G on S is a circular drawing of G .

The definition of intersection compatible coloring extends to boundary drawings in a natural way. If G is a graph with a boundary drawing f on a surface S , a coloring c of G is intersection compatible if for every pair of edges $e = (u, v)$ and $e' = (u', v')$, if the paths $f(e)$ and $f(e')$ intersect, then the two edges have at least one color in common, i.e., $\{c(u), c(v)\} \cap \{c(u'), c(v')\} \neq \emptyset$. We will need the following simple result describing how the Euler characteristic of a surface changes when we cut this surface along a path.

Lemma 12. *Let S be a surface with a boundary, and p be a simple path on S between two boundary points. Cutting the surface S along the path p yields a surface with Euler characteristic increased by 1.*

Proof. Let S be a surface with boundary, and h the Euler characteristic of S . Let p be a path between two boundary points of S , and S' be the surface obtained by cutting S along the path p . Let T be a triangulation of S . Without loss of generality, we can assume that the path p is along the edges of the triangulation T . Consider the triangulation T' where each edge and each vertex of the path p were duplicated as in Figure 4. Let n be the number of vertices in the path p . Then, T' is a triangulation for S' , and the Euler characteristic of S' is:

$$|V(T')| - |E(T')| + |F(T')| = (|V(T)| + n) - (|E(T)| + n - 1) + |F(T)| = h + 1$$

□

Note that in some cases, cutting a surface along a path can disconnect the surface, or increase the number of boundary components as in Figure 4. To prove the theorem, the key ingredient is to prove an extension of Lemma 9 for graphs with a boundary drawing on an arbitrary surface. This is done in the following lemma.

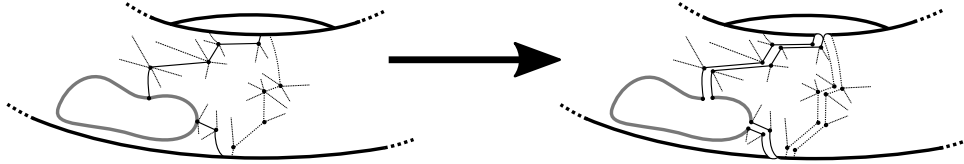


Figure 4: Cutting a surface (here a torus with a boundary) along a path. The bold grey line is the boundary of the surface.

Lemma 13. *There exists a constant K_0 such that, if G is a graph with n vertices with a boundary drawing on a surface S with Euler characteristic h , and c is a Grundy k -coloring of G which is also intersection compatible, then $k \leq K_0 \log n + 2(1 - h)$.*

Proof. We show the result by induction on h , the Euler characteristic of the surface S . If $h = 1$, then the surface S is a disk, and the result follows from Lemma 9. Suppose that $h < 1$, and let G be a graph with a boundary drawing f on a surface S with Euler characteristic h . Let c be a Grundy coloring of G that is intersection compatible for this drawing. We distinguish two cases:

Case 1. There is an edge e of G such that cutting the surface S along the path $f(e)$ leaves the surface connected. Let a and b be the two colors of the endpoints of this edge. We consider S' , the surface obtained by cutting S along $f(e)$, and G' the graph induced by the vertices with a color different from a or b . Then, the boundary drawing of G on S induces a boundary drawing of G' on S' . Indeed, if (u, v) is an edge of G' , then the path corresponding to (u, v) cannot cross $f(e)$ since otherwise one of the endpoints u or v would be colored either a or b since the coloring is intersection compatible.

Moreover, by Lemma 12, the surface S' has Euler characteristic $h + 1$. Using the induction hypothesis on G' , with the coloring induced by c on G' , we know that c colors G' with at most $K_0 \log n + 2(1 - (h + 1))$ colors. Since the vertices we removed were the ones colored a or b , this implies that the coloring c of G uses at most $K_0 \log n + 2(1 - h)$ colors, which proves the induction step.

Case 2. For any edge e , cutting S along the path $f(e)$ disconnects the surface. In this case, we claim that the graph G has a boundary drawing on a disk such that c is intersection compatible for this drawing. To prove this, we first remark that we can assume that there is only one boundary component on the surface S . Indeed, suppose that this is not the case. Consider the boundary that contains a vertex v with color k , the largest color used by c . Let G' be the graph induced by all the vertices on this boundary. Then all the colors are present in G' . Indeed, G' contains the vertex v with color k . Since c is a Grundy coloring, there are vertices v_1, \dots, v_{k-1} adjacent to v , with colors $1, \dots, k - 1$ respectively. For all i , the vertex v_i is on the same boundary as v since otherwise cutting along the path corresponding to the edge (v, v_i) would leave the surface connected. Let c' be the coloring induced by c on G' . If the result holds for G' with coloring c' , it also holds for G with coloring c since c and c' use the same number of colors. So we now assume that S has one unique boundary.

Let u_1, \dots, u_n be the vertices of G in the order they appear on the boundary of S , and consider the circular drawing of G using the same ordering u_1, \dots, u_n around the circle. To prove the result, we only need to show that c is also intersection compatible for this new drawing. Indeed, if this is the case, then the result immediately follows from Lemma 9. Let e_1 and e_2 be two edges of G that intersect in the circular drawing of G . Without loss of generality, we can assume that $e_1 = (u_1, u_i)$, and $e_2 = (u_j, u_l)$, with $1 < j < i$, and $i < l \leq n$. Let p_1 and p_2 be the paths on S corresponding to these two edges. Assume by contradiction that in the drawing of G on S , the two paths p_1 and p_2 do not intersect, we will show that cutting along p_1 does not disconnect the surface S . For this, it is enough to show that there is a path not intersecting p_1 going from one side of p_1 to the other side. We will construct a path from u_{i-1} to u_{i+1} that does not intersect p_1 . The path is as follows. Go from u_{i-1} to u_j by following the border of S , then follow the path p_2 from u_j to u_l , and finally go from u_l to u_{i+1} by following again the border of S . This path does not intersect p_1 . Hence cutting along p_1 leaves the surface S connected, a contradiction.

Thus, we know that for every pair of edges e_1 and e_2 , if the segments corresponding to e_1 and e_2 in the circular drawing of G intersect, then the paths corresponding to these two edges on the drawing of

G on S also intersect. Consequently, c is intersection compatible for this circular drawing of G , and the result follows from Lemma 9. □

We now have everything we need to prove the theorem.

(*proof of Theorem 11*). The proof follows the same argument as the proof of Theorem 7. Let G be a graph with Euler characteristic h , v a vertex of G , and c a coloring of $G - v$. We want to show that after applying procedure **GreedyRecolor**, the neighborhood of v is colored using at most $K_0 \log \Delta + 2(2 - h)$ colors. Let S be a closed surface with Euler characteristic h such that there is an embedding of G on S . Let G' be the graph whose vertices are the neighbors of v , and such that there is an edge between u_1 and u_2 if and only if $c(u_1) \neq c(u_2)$ and there is a bichromatic path from u_1 to u_2 . Without loss of generality, we can assume that the colors are ordered by decreasing size: $\text{size}_c(1) \geq \text{size}_c(2) \geq \dots$. Then, as before, c is a Grundy coloring of G' . Indeed, if there is a vertex v and a color $i < c(v)$ such that v has no neighbor colored i , then the bichromatic exchange $\langle v, i \rangle$ increases inequality.

Moreover, if S' is the surface S where a small disk around vertex v was removed, then from the embedding of G on S , we can construct a boundary drawing of G' on S' . The coloring c induces an intersection compatible drawing of G' for this drawing. Since the Euler characteristic of S' is $h - 1$, by Lemma 13, at most $K_0 \log \Delta + 2(2 - h)$ are used by c on G' . Consequently, at least one color is not present in the neighborhood of v . □

6 Counter example

The goal of this section is to exhibit concrete example of graphs with small chromatic number that need a large number of colors to be colored by an online algorithm using bichromatic exchanges. Unfortunately, this example does not work for general algorithms with bichromatic exchanges, but only to more special cases with the following restrictions on the algorithm:

- The algorithm is specified as input the number k of colors it is allowed to use to color the graph.
- If one of the k colors is available, the algorithm must select it, without performing any bichromatic exchange.
- The algorithm always chooses the smallest color available.

These assumptions might seem a bit restrictive, but all the algorithms we described above satisfy these constraints. Getting a counter example that would work in the general case is more complicated as it is difficult to quantify how much the graph can be recolored by bichromatic exchanges performed by the algorithm. We show the following theorem:

Theorem 14. *There is a sequence of graph $(G_k)_{k \geq 0}$ such that for all $k \geq 4$, G_k is 3 colorable, and no online algorithm with bichromatic exchanges with the restrictions above can color G_k with k colors or less.*

Proof. Let A be an online algorithm with bichromatic exchanges with the restrictions above. First observe that as long as one color is available, the algorithm A behaves exactly as FIRST-FIT: it assigns the smallest color available without changing the color of previously colored vertices. We denote by T_k the fibonacci tree with the following definition. The tree T_1 is a single vertex, and for $k > 0$, T_k is composed of a root vertex u , to which we attach the trees T_1, \dots, T_{k-1} .

It was proved in [Bea76], that if the vertices of T_k are presented starting from the leaves, and going up to the root, then FIRST-FIT needs k colors to color T_k , and the root of T_k is colored with color k . We build the graph H_k in the following way:

- for each $1 \leq i \leq k$, we add a copy of T_i , with vertex u_i as the root,
- for each $i < j$, and each $i' \neq i$ and $j' \neq j$ with $i' \neq j'$, we add one copy of $T_{i'}$ and one copy of $T_{j'}$ with roots $w_{i,j,i',j'}^1$ and $w_{i,j,i',j'}^2$ respectively,
- we add all the following edges: $(u_i, w_{i,j,i',j'}^1)$, $(u_j, w_{i,j,i',j'}^2)$, and $(w_{i,j,i',j'}^1, w_{i,j,i',j'}^2)$,

- finally, a vertex u is added, with u adjacent to all the u_i .

The graph H_k is presented starting from the leaves of all the copies of T_i , and going up to the roots. The vertex u is presented last. Since H_k is 2-degenerate, it is 3 colorable. We prove that the algorithm A can't color H_k if it is given exactly k colors. Indeed, since the algorithm always color immediately a vertex if one color is available, it will be able to color all the vertices except for u , by applying the rules of FIRST-FIT. When it tries to color u , all the colors are already present in the neighborhood of u , so the algorithm might try to remove one color using local bichromatic exchanges. However, we will show that at this point, no local bichromatic exchange can remove one color from the neighborhood of u . More precisely, we will show that any local bichromatic exchange will preserve the following invariants:

1. The vertices u_i all have different colors.
2. For every indices i and j , with $i \neq j$, and every pair of colors $a \neq b$ with $a \neq c(u_i)$, and $b \neq c(u_j)$, there is a path on four vertices (u_i, w^1, w^2, u_j) such that $c(w_1) = a$ and $c(w_2) = b$.

By construction, these invariants are satisfied before any bichromatic exchange is made. Suppose now that there is a coloring c that satisfies these invariants. Consider the bichromatic exchange $\langle u_i, x \rangle$, and let $j \neq i$ such that $x = c(u_j)$. We will show that the coloring $c' = \langle u_i, x \rangle(c)$ still satisfy the two invariants above.

Using the fact that c satisfies the second invariant, we know that there is a bichromatic path from u_i to u_j . Consequently, applying the bichromatic exchange $\langle u_i, x \rangle$, swaps the colors of the vertices u_i and u_j . Consequently, c' still satisfies the invariant 1. We only need to show that c' also satisfies the second invariant. Consider two indices i' and j' , and two colors $a \neq c'(i')$ and $b \neq c'(j')$. We consider the following three cases:

- If $i' \neq i$ and $j' \neq j$, then $c'(u_{i'}) = c(u_{i'})$ and $c'(u_{j'}) = c(u_{j'})$. Since c satisfies the second property, there are two vertices w^1 and w^2 such that $c(w^1) = a$ and $c(w^2) = b$ such that $(u_{i'}, w^1, w^2, u_{j'})$ is a path on four vertices. Since the colors of w^1 and w^2 also do not change during the bichromatic exchange, we also have $c'(w^1) = a$ and $c'(w^2) = b$.
- If $i' = i$ and $j' \neq j$, then $c'(u_i) = c(u_j)$, and $c'(u_{j'}) = c(u_{j'})$. If $a \neq c(u_j)$, then in the coloring c there was a path on four vertices with successive colors $(c(u_i), a, b, c(u_{j'}))$. This path now has colors $(c(u_j), a, b, c(u_{j'}))$. If $a = c(u_i)$ and $b \neq c(u_j)$, then with the coloring c there was a path from u_i to $u_{j'}$ with colors $(c(u_i), c(u_j), b, c(u_{j'}))$. This path is now colored $(c(u_j), a = c(u_i), b, c(u_{j'}))$. Finally, if $a = c(u_i)$ and $b = c(u_j)$, the path with colors $(c(u_i), c(u_j), c(u_i), c(u_{j'}))$ now has the necessary colors.

The argument is symmetrical if $i' \neq i$ and $j = j'$.

- If $i' = i$ and $j' = j$, the same argument as above can be used to prove that there is still a path between u_i and u_j with colors $(c'(u_i), a, b, c'(u_j))$. The idea is that colors different than $c(u_i)$ and $c(u_j)$ are unchanged, and vertices colored $c(u_i)$ or $c(u_j)$ gets their color swapped.

Consequently, the invariant is preserved when we perform any local bichromatic exchange. Consequently, the algorithm A won't be able to color H_k if it is given exactly k colors. However, the argument relies on the fact that the algorithm does not perform bichromatic exchanges before considering the vertex u . Consequently, the algorithm might still manage to color H_k if it is given less than k colors. This issue can be solved by considering $G_k = \bigcup_{i \leq k} H_i$. For any $k' \leq k$, the algorithm A will fail to color H'_k , and consequently G_k , and the theorem holds. \square

7 Conclusion

We studied a variation of online algorithms where the algorithm is allowed to recolor some of the previously colored vertices. We have shown algorithms for several classes of graphs. In the case of planar graphs, it remains open whether the $O(\log \Delta)$ bound can be improved or not.

We introduced the notion of intersection compatible coloring which could be interesting to look at on its own. In particular, a relevant question could be to try to characterize the graphs which have a drawing on a given surface S such that any two intersecting edges share a common color.

Finally, from a reconfiguration point of view, we studied a recoloring problem with some locality constraints on the transformations that we are allowed to make. It could be interesting to look at standard reconfiguration problems when we add these constraints. An example of problems to look at could be, deciding if a given coloring can be transformed into an other using local bichromatic exchanges, or given two vertices u and v , deciding if there is a sequence of bichromatic exchanges local to u that change the color of v .

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